On proving the termination of algorithms by machine

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Abstract

Proving the termination of a recursively defined algorithm requires a certain creativity of the (human or automated) reasoner for inventing a hypothesis whose truth implies that the algorithm terminates. We present a reasoning method for simulating this kind of creativity by machine. The proposed method works automatically, i.e. without any human support. We show, (1) how a termination hypothesis for an algorithm is synthesized by machine, (2) which knowledge about algorithms is required for an automated synthesis, and (3) how this knowledge is computed. Our method solves the problem for a relevant class of algorithms, including classical sorting algorithms and algorithms for standard arithmetical operations, which are given in a pure functional notation. The soundness of the method is proved and several examples are presented for illustrating the performance of the proposal. The method has been implemented and proved successful in practice.

1. Introduction

A central problem in the development of correct software is to verify that algorithms always terminate, provided the intended operations have decidable domains. Non-terminating algorithms compute partial operations, hence machine resources are wasted if a given input is not in the domain of the implemented operation. Also manpower is wasted with the debugging of those algorithms and the frustration caused by non-terminating programs is a common experience of programmers and computer scientists. Therefore techniques to verify termination

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are of considerable interest, but since the halting problem is the "classical" undecidable problem there is no procedure which proves or disproves the termination of all algorithms.

If we have an idea why an algorithm terminates, we can use an automated theorem prover for verification. But finding the right argument which implies termination usually requires some creativity and it is our aim to simulate this kind of creativity by machine. Hence we are concerned here essentially with theorem discovery rather than with theorem proving.

The termination of a recursively defined algorithm is proved by invention of a well-founded order relation for the algorithm and subsequent verification that the arguments in each recursive call are smaller—in the sense of the invented order—than the initially given input. The basic approach is to invent a termination function, also called convergence function or bound function, which decreases the arguments in each recursive call in the sense of some known well-founded relation. However, finding a successful termination function is the crucial step in this approach and this is the creativity we intend to mechanize.

This paper is concerned with the automation of termination proofs for a certain class of algorithms. We consider only algorithms which terminate according to well-founded relations based on the so-called size order. This order compares data objects by their size, e.g. stacks are compared by their depth, lists by their length, trees by the number of their nodes, etc. Suppose, for instance, that we want to prove the termination of a sorting algorithm for lists of natural numbers. This data structure can be defined in our notation by

```plaintext
structure empty add(head : number tail : list) : list ,
```

where number (standing for the natural numbers) denotes another data structure defined elsewhere. The symbols empty and add are the constructors of list, i.e. each list equals empty or else can be constructed by applying add to elements of the data structures number and list. The symbols head and tail are the selectors of add and serve as kinds of inverse operations to the constructor, yielding the first element of a non-empty list and the list with the first element removed respectively.

As an example of an algorithm in our notation we define an algorithm remove which eliminates all occurrences of a number n from a list x:

```plaintext
function remove(n : number x : list) : list \equiv 
  if x = empty then x
  if x \neq empty \land head(x) = n then remove(n tail(x))
  if x \neq empty \land head(x) \neq n then add(head(x) remove(n tail(x))) .
```

Given another algorithm function minimum(x : list) : number \equiv \cdots which returns a minimal element of a non-empty list x (in the sense of some order relation for number defined elsewhere), we define a sorting algorithm for list by
function sort(x : list) : list  \leftarrow
if x = empty then empty
if x \neq empty then
  add(minimum(x) sort(remove(minimum(x) x))). \textsuperscript{1}
\]

To verify the termination of \textit{sort} a typical argumentation would read: "We have to find a well-founded relation \(<_\text{\texttt{\#}}\) such that \(\text{remove}(\text{minimum}(x) x) \prec_{\text{\texttt{\#}}} x\) whenever \(x \neq \text{empty}\) holds. Our first observation is that \(\text{remove}(n x)\) returns \(x\) or a \textit{list} which is shorter than \(x\), in symbols
\[
\text{remove}(n x) \prec_{\text{\texttt{\#}}} x, \quad (1.1)
\]

hence \(<_\text{\texttt{\#}}\) (standing for "shorter than") could be the well-founded relation we are looking for. Now searching for a condition which implies \(\text{remove}(n x) \prec_{\text{\texttt{\#}}} x\) we find that this requirement is satisfied iff \(n\) is a member of \(x\). So if we define \textit{list}-membership by
\[
function \text{member}(n : \text{number} x : \text{list}) : \text{bool} \leftarrow
if x = empty then false
if x \neq empty \land \text{head}(x) = n then true
if x \neq empty \land \text{head}(x) \neq n then \text{member}(n \text{tail}(x)), \quad (1.2)
\]

we can prove that
\[
\forall n : \text{number} \forall x : \text{list} \quad \text{member}(n x) \leftrightarrow \text{remove}(n x) \prec_{\text{\texttt{\#}}} x. \quad (1.3)
\]

Our next observation is that each non-empty \textit{list} contains its minimum, i.e. we can also prove
\[
\forall x : \text{list} \quad x \neq \text{empty} \rightarrow \text{member}(\text{minimum}(x) x). \quad (1.4)
\]

From (1.4) and (1.3) we obtain finally
\[
\forall x : \text{list} \quad x \neq \text{empty} \rightarrow \text{remove}(\text{minimum}(x) x) \prec_{\text{\texttt{\#}}} x \quad (1.5)
\]

and the termination of \textit{sort} is proved".\textsuperscript{2}

Our termination proof procedure mechanizes this chain of reasoning in the spirit of the generate-and-test paradigm: Analyzing algorithms such as \textit{remove}, a machine \textit{recognizes} properties such as (1.1) and \textit{synthesizes} algorithms as (1.2) such that properties as (1.3) hold. Then \textit{termination hypotheses} such as (1.4) are

\textsuperscript{1}This algorithm also "purges" a list, i.e. multiple occurrences of list elements are eliminated. We could have defined a "real" sorting algorithm using \textit{delete} \textsuperscript{2} instead of \textit{remove} but we prefer this version of \textit{sort} because we need \textit{remove} for subsequent illustrations. It is easily seen that our technique works for \textit{delete} in the same way as it does for \textit{remove}.

\textsuperscript{2}One may observe that the definition of \textit{minimum} is irrelevant for the termination of \textit{sort} as long as formula (1.4) holds. Hence \textit{sort} also terminates if e.g. \textit{minimum} returns the last element of a non-empty \textit{list} (in which case \textit{sort} reverses a purged list), or if \textit{minimum}(x) returns \textit{head}(x) (in which case \textit{sort} computes the purged list).
computed which are sufficient for the termination of an algorithm such as \textit{sort}. Finally an \textit{induction theorem prover} is used as a "tester" for verifying the generated termination hypotheses and this completes the termination proof. Since we do not need special theorem proving abilities we only discuss the "generate" mode of our technique and assume the availability of an induction theorem prover.

The remainder of this paper is organized as follows: The environment in which we define and use data structures and algorithms is settled in Section 2 and termination of algorithms is discussed in Section 3. Section 4 shows how we prove inequalities by the technique of estimation: We define the \textit{estimation calculus}, i.e. a calculus which provides a deductive notion for the (semantical) size order. There we assume the availability of certain knowledge about algorithms which is relevant for a termination proof, such as (1.1), (1.2) and (1.3) in the example above. Using the estimation calculus, termination hypotheses for algorithms such as (1.4) can be generated by machine as shown in Section 5. Section 6 is concerned with an automatic acquisition of the knowledge which is necessary for the generation of termination hypotheses: We present a method for recognizing whether an algorithm is \textit{argument-bounded}, i.e. whether the size of an algorithm's result is always smaller than or equal to the size of one of its input arguments. With this technique, for instance, \textit{remove} is recognized as argument-bounded and a fact such as (1.1) is found by machine. Then we show how a so-called \textit{difference algorithm} can be synthesized for each argument-bounded algorithm. Such an algorithm returns \textit{true} for a list of input arguments, if and only if the given argument-bounded algorithm (applied to the same input arguments) returns a result \textit{strictly} smaller than one of its arguments. We also show how each difference algorithm is optimized. For our example \textit{member} is synthesized as the optimized difference algorithm for \textit{remove} such that a property like (1.3) holds. Shortcomings of our proposal are discussed in Section 7, and Section 8 presents related work. A collection of sorting algorithms whose termination has been automatically proved by an implementation of the method is given in Appendix A.

2. Formal preliminaries

2.1. Data structures and algorithms

Suppose we have a collection of data structures and algorithms operating on them. For instance, we may have a data structure \textit{number} given as

\begin{equation}
\text{structure } 0 \text{ succ( pred : number) : number} \quad (2.1)
\end{equation}

that denotes the set of natural numbers with successor function \textit{succ}, predecessor function \textit{pred} and 0 for zero, and we may have an algorithm

\begin{equation}
\begin{align*}
\text{function } & \text{plus}(x, y : \text{number}) : \text{number} \leftarrow \\
& \text{if } x = 0 \text{ then } y \\
& \text{if } x \neq 0 \text{ then } \text{succ}(\text{plus}(\text{pred}(x) y)),
\end{align*} \quad (2.2)
\end{equation}
which computes addition. The domains our algorithms are operating on are given as sets of constructor ground terms, e.g. the data structure $\text{number}$ defines a set $\mathcal{T}(\Sigma)_{\text{number}} = \{0, \text{succ}(0), \text{succ}(\text{succ}(0)), \ldots\}$ of constructor ground terms of sort $\text{number}$ to represent the set of natural numbers. This means that 0 and $\text{succ}$ are used as the constructors of the data structure $\text{number}$. Each argument position of a constructor is associated with a so-called selector (or destructor). Here we have $\text{pred}$ as the (only) selector of $\text{succ}$, i.e. $\text{pred}$ denotes the function which maps a number to a number such that $\text{pred}(\text{succ}(q)) = q$ is satisfied. As another example we have already seen the data structure $\text{list}$

$$\text{structure empty add(head : number tail : list) : list}, \quad (2.3)$$

which defines the set $\mathcal{T}(\Sigma)_{\text{list}} = \{\text{empty}, \text{add}(0 \text{ empty}), \text{add}(\text{succ}(0) \text{ empty}), \text{add}(0 \text{ add}(\text{succ}(0) \text{ empty})), \ldots\}$ of constructor ground terms of sort $\text{list}$ such that $\text{head}(\text{add}(n k)) = n$ and $\text{tail}(\text{add}(n k)) = k$ is satisfied. We may also define a data structure $\text{bool}$ by

$$\text{structure true false : bool}, \quad (2.4)$$

to define predicates like, for instance, the “less-than” order on numbers by

$$\text{function lt(x, y : number) : bool} \iff$$

$$\text{if } y = 0 \text{ then false}$$

$$\text{if } y \neq 0 \land x = 0 \text{ then true}$$

$$\text{if } y \neq 0 \land x \neq 0 \text{ then } \text{lt}(\text{pred}(x) \text{ pred}(y)). \quad (2.5)$$

Subsequently we write boolean-valued functions like predicates, e.g. we use $\text{lt}(x y)$ as an abbreviation for $\text{lt}(x y) = \text{true}$ and $\neg \text{lt}(x y)$ abbreviates $\text{lt}(x y) = \text{false}$.

We always demand that an algorithm $f$ is well-formed, i.e. each function symbol used in the definition of $f$ either is introduced before by a data structure or another algorithm or otherwise is the new symbol $f$ in which case the algorithm is recursively defined. Now assume that the conditions of the cases in an algorithm exclude each other, as for instance the conditions of the algorithm $\text{lt}$, viz. $y = 0$, $y \neq 0 \land x = 0$ and $y \neq 0 \land x \neq 0$, do. We call these algorithms deterministic and by this requirement we can think of an interpreter $I$ for the algorithms which implements the language in the following way: For each ground term given as input (i.e. a term without variable symbols), the interpreter evaluates the term by applying the algorithms which are “called” in the term to the evaluation of their arguments until no further applications are possible. If, for instance, $\text{plus}(\text{succ}(0) \text{ succ}(0))$ is provided as input, $\text{plus}$ is applied to $(\text{succ}(0) \text{ succ}(0))$ yielding $\text{succ}(\text{plus}(0 \text{ succ}(0)))$, then $\text{plus}$ is applied to $(0 \text{ succ}(0))$ yielding $\text{succ}(0)$, and therefore $\text{succ}(\text{succ}(0))$ is the final result of the evaluation.

Next assume that all algorithms in the collection terminate, as $\text{plus}$ and $\text{lt}$ do, and therefore the evaluation of each ground term halts after finitely many steps with a ground term as result. Then we may view the interpreter $I$ as a total mapping $I: \mathcal{T}(\Sigma) \rightarrow \mathcal{T}(\Sigma)$ from the set $\mathcal{T}(\Sigma)$ of ground terms (built with the function symbols $\Sigma$ introduced by the collection) to itself.
Finally assume that each algorithm is case-complete, i.e. the case analysis of an algorithm is complete in the sense that for each n-tuple of constructor ground terms given as an input to an algorithm with n arguments, at least one of the case conditions is satisfied for this input. The algorithm is case-complete, for instance, because either \( y = 0 \), \( y \neq 0 \land x = 0 \) or \( y \neq 0 \land x \neq 0 \) is true for each pair of constructor terms of sort number substituted for \( x \) and \( y \). Now we may view the interpreter \( I \) as a total mapping \( I : \mathcal{T}(\Sigma) \rightarrow \mathcal{T}(\Sigma^c) \) from the set \( \mathcal{T}(\Sigma) \) of ground terms to the set \( \mathcal{T}(\Sigma^c) \) of constructor ground terms. This is because by termination each evaluation ends with a non-recursive case, by case-completeness the condition of this case is satisfied and well-formedness then guarantees that a constructor ground term is returned.

We may also evaluate terms containing variables by using variable bindings of form \([x^*/q^*]\), where \( x^* \) is a list of distinct variables and \( q^* \) is a list of constructor ground terms. We write, for instance, \( I[n, m/succ(0), 0] \) (plus(n succ(m))) = succ(succ(0)) to indicate that the interpreter \( I \) evaluates plus(n succ(m)) under the variable binding \([n, m/succ(0), 0]\) to succ(succ(0)).

### 2.2. Admissible specifications and theories

Statements about the data structures and algorithms can be formulated as formulas of a many-sorted first-order language \( \mathcal{F}(\Sigma, \mathcal{V}) \). The names of the data structures, e.g. number, list and bool, are used as sort symbols of our many-sorted language and we assume a set of variable symbols \( \mathcal{V}_s \) for each sort symbol \( s \). The function symbols introduced by the data structures and algorithms, e.g. succ, pred, plus, it, ..., define the signature \( \Sigma \) from which the set of terms \( \mathcal{T}(\Sigma, \mathcal{V}) \) is constructed, where \( \mathcal{V} \) is the set of all variable symbols. The only predicate symbols are \( \text{TRUE} \) and \( \text{FALSE} \) (for truth and falsity) and \( = \) which is a binary predicate denoting equality. Using \( \text{TRUE} \), \( \text{FALSE} \) and equations of the form \( t_1 = t_2 \) as atomic formulas, connectives like \( \rightarrow, \land, \lor \), etc. and the quantifiers \( \forall \) and \( \exists \), the set \( \mathcal{F}(\Sigma, \mathcal{V}) \) of many-sorted first-order formulas is defined as usual. We write \( x \in \mathcal{V}_s \) to denote that \( x \) is a variable of sort \( s \), and if \( w \) is a finite and non-empty sequence of sort symbols, i.e. \( w = s_1 \ldots s_n \), then \( \mathcal{V}_w \) stands for the cartesian product \( \mathcal{V}_{s_1} \times \cdots \times \mathcal{V}_{s_n} \) and \( \mathcal{T}(\Sigma^c)_w \) abbreviates \( \mathcal{T}(\Sigma^c)_{s_1} \times \cdots \times \mathcal{T}(\Sigma^c)_{s_n} \). If we write "\( x^* \in \mathcal{V}_w \)", it is implicitly assumed that \( x^* \) consists of distinct variables.

A collection of data structures and well-formed algorithms is called an admissible specification iff each algorithm in the collection terminates, is deterministic and is case-complete. So for each admissible specification there is exactly one interpreter \( I \) such that each ground term \( t \) is evaluated by \( I \) to a constructor ground term \( I(t) \) as the result of the evaluation and we may use \( I \) for defining an interpretation \( \mathcal{I} \) for the formulas in \( \mathcal{F}(\Sigma, \mathcal{V}) \).

The universe of \( \mathcal{I} \) is the set \( \mathcal{T}(\Sigma^c) \) of constructor ground terms and \( \mathcal{I} \) assigns each function symbol \( f \in \Sigma \) the mapping computed by the interpreter \( I \). We write \( \mathcal{I} \models \varphi \) if \( \mathcal{I} \) satisfies the formula \( \varphi \) where \( \mathcal{I} \models \forall x^*: w \varphi \) iff \( \mathcal{I}[x^*/q^*] \models \varphi \) for all \( n \)-tuples of constructor ground terms \( q^* \in \mathcal{T}(\Sigma^c)_w \) which can be substituted for
the variables in $x^* \in V_w$. The meaning of TRUE and FALSE, of the connectives and of the equality sign is as usual. $\mathcal{I}$ is a model of a formula $\varphi$ (or a set of formulas $\Phi$) iff $\mathcal{I} \models \varphi$ (or $\mathcal{I} \models \Phi$, i.e. $\mathcal{I} \models \varphi$ for all $\varphi \in \Phi$). We call $\mathcal{I}$ a standard interpretation for $S$ iff $\mathcal{I}$ is obtained from the interpreter of an admissible specification $S$ as just defined.

A statement $\varphi$ about the data structures and algorithms in an admissible specification is true iff $\mathcal{I} \models \varphi$. The set of all true statements is given by the theory $Th(\mathcal{I})$ of $\mathcal{I}$, i.e. the set of all closed first-order formulas $\varphi$ such that $\mathcal{I} \models \varphi$. Since there is only one standard interpretation $\mathcal{I}$ for an admissible specification $S$ we may also associate the specification $S$ with a theory $Th(S) := Th(\mathcal{I})$. The theory $Th(S)$ is the set of all true first-order formulas about the expressions in $S$ and we say that an admissible specification specifies a theory. For instance, if $S$ consists only of the data structure number and the algorithm plus, then $Th(S)$ is the set of all true statements about $0$, pred, succ and plus, e.g. commutativity and associativity of plus, cancellation properties of plus, etc. We therefore write

$$\left[ \forall x, y : \text{number } \text{plus}(x \ y) = \text{plus}(y \ x) \right] \in Th(S) ,$$

$$\left[ \forall x, y, z : \text{number } \text{plus}(x \ \text{plus}(y \ z)) = \text{plus}(\text{plus}(x \ y) \ z) \right] \in Th(S) ,$$

but

$$\left[ \forall x : \text{number } x = \text{plus}(x \ x) \right] \not\in Th(S) ,$$

etc.

An admissible specification $S_{n+1}$ can be obtained by extending an admissible specification $S_n$ with a new data structure or a new algorithm, where we always start with the initial (admissible) specification $S_0$ which consists only of the data structure bool. On extending $S_n$ by a new data structure only trivial syntactic features have to be tested, e.g. that the function symbols for the constructors and selectors are "new" with respect to the symbols already used in $S_n$. On extension by an algorithm, well-formedness, case-completeness, determinism and termination have to be verified. The test for well-formedness is trivial. The tests for case-completeness and determinism can be avoided by an if–then–else conditional, or otherwise an induction theorem prover (cf. Section 2.3) is used for verification. The termination test is the subject of the subsequent developments.

2.3. Theorem proving

To verify a statement $\varphi$ about the expressions in an admissible specification, i.e. to test whether $\varphi \in Th(S)$, some deduction is needed. Therefore each data structure and each algorithm of the specification is translated into a set of formulas yielding the set of axioms $\Phi(S)$ of the specification $S$. Then a (sound) first-order calculus is used to infer the statement $\varphi$ from the axioms $\Phi(S)$.

The axioms are obtained by a uniform translation process. Each data structure $s$ of the specification is translated into a set $REP_s$ of representation formulas for $s$
which define properties of the constructors and selectors. For instance, the data structure list yields the axioms

\begin{align}
\forall k : \text{list } \forall n : \text{number} & \quad \text{empty} \neq \text{add}(n \; k), \quad (2.6) \\
\forall k_1, k_2 : \text{list } \forall n_1, n_2 : \text{number} & \quad \text{add}(n_1 \; k_1) = \text{add}(n_2 \; k_2) \rightarrow n_1 = n_2 \land k_1 = k_2, \quad (2.7) \\
\forall k : \text{list } & \quad k = \text{empty} \vee k = \text{add}(\text{head}(k) \; \text{tail}(k)), \quad (2.8) \\
\forall k : \text{list } \forall n : \text{number} & \quad \text{head}(\text{add}(n \; k)) = n, \quad (2.9) \\
\forall k : \text{list } \forall n : \text{number} & \quad \text{tail}(\text{add}(n \; k)) = k. \quad (2.10)
\end{align}

However, using only (2.6)–(2.10) nothing is said about what will happen if head and tail are applied to empty. Since selectors essentially denote partial mappings defined only on arguments built with their constructors, we are free to associate head(empty) (and tail(empty)) with any constructor ground term of sort number (and list respectively). But we also need to make such assignments because otherwise head(empty) and tail(empty) cannot be evaluated to constructor ground terms. For avoiding this situation we associate with each data structure s a so-called witness term Vs. The witness term Vs is obtained as the application of the leftmost irreflexive constructor in the definition of the data structure s to the witness terms of its argument sorts.3 Using this notion, we define sel(cons(...)) = Vs for each selector sel and each constructor cons such that sel does not belong to the constructor cons. Consequently we demand that each data structure s provides one irreflexive constructor at least. But this must be demanded anyway because otherwise the set \( \mathcal{F}(\Sigma^s) \) of constructor ground terms of sort s would be empty. Here we have 0 as witness term of number, i.e. \( \forall \text{number} = 0 \), and consequently

\begin{equation}
\text{head}(\text{empty}) = 0 \quad (2.11)
\end{equation}

is obtained. We also find \( \forall \text{list} = \text{empty} \) and therefore

\begin{equation}
\text{tail}(\text{empty}) = \text{empty}. \quad (2.12)
\end{equation}

The formulas (2.6)–(2.12) constitute the set REPlist of representation formulas for list, and we define \( \text{REP}_\text{list} \subseteq \Phi(S) \).

Also algorithms are translated into axioms: Each algorithm

\footnote{A function symbol is called irreflexive iff its rangesort is different from all its domainsorts, and otherwise it is called reflexive. For instance, each constant symbol, e.g. 0 or empty, is irreflexive (simply because there are no domainsorts for it) and head is an irreflexive (non-constant) function symbol. The function symbols pred, succ, add, tail and plus all are reflexive. The choice of the left-most irreflexive constructor is completely arbitrary, i.e. the rightmost irreflexive constructor or any other would do as well.}
function $f(x^*; w) : s \iff$

if $\varphi_1$ then $r_1$

: 

if $\varphi_k$ then $r_k$

(2.13)

is translated into a set

$$\text{DEF}_f = \{ \forall x^*: w \, \varphi_1 \rightarrow f(x^*) = r_1, \ldots, \forall x^*: w \, \varphi_k \rightarrow f(x^*) = r_k \}$$

(2.14)

of definition formulas for $f$, and we demand $\text{DEF}_f \subseteq \Phi(S)$ for each algorithm $f$ in the specification $S$. For instance, the algorithm $\text{plus}$ is translated into the set of definition formulas

$$\text{DEF}_{\text{plus}} = \{ \forall x, y : \text{number} \, x = 0 \rightarrow \text{plus}(x, y) = y, \forall x, y : \text{number} \, x \neq 0 \rightarrow \text{plus}(x, y) = \text{succ}(\text{plus}(\text{pred}(x), y)) \}. \quad (2.15)$$

It can be verified that all axioms obtained from the data structures and algorithms are satisfied by the standard interpretation, i.e. $\mathcal{I} \models \Phi(S)$, and therefore $\Phi(S) \subseteq \text{Th}(S)$, cf. [22]. Hence $\mathcal{I}$ is called the standard model of the axioms $\Phi(S)$ and it is also called the standard model of the specification $S$.

Our termination proof procedure generates formulas $\varphi$ from an algorithm $f$ such that $\varphi \in \text{Th}(S)$ entails the termination of $f$. Therefore some machine support for testing $\varphi \in \text{Th}(S)$ is required if termination of algorithms has to be verified by machine. This support is provided by an induction theorem proving system (see e.g. [24] for a survey and further references). Roughly speaking, such a system tries to verify $\Phi(S) \cup I(S) \vdash \varphi$ where $\vdash$ denotes derivability in some sound first-order calculus $K$ and $I(S) \subseteq \text{Th}(S)$ is some decidable set of first-order formulas associated with $S$, called the induction axioms of $S$. If successful, then $\varphi \in \text{Th}(S)$ because all models of $\Phi(S) \cup I(S)$ are also models of $\varphi$ (with $K$ being sound), and then in particular $\mathcal{I} \models \varphi$ because $\mathcal{I} \models \Phi(S) \cup I(S)$ (with $\Phi(S) \cup I(S) \subseteq \text{Th}(S)$). Consequently $\Phi(S) \cup I(S) \vdash \varphi$ is sufficient for $\varphi \in \text{Th}(S)$ to hold. So for an implementation of our method each (semantical) requirement of the form “$\varphi \in \text{Th}(S)$” which is stipulated subsequently must be replaced by a call of an induction theorem prover to test for “$\Phi(S) \cup I(S) \vdash \varphi$”.

3. Terminating algorithms

Let function $f(x_1 : s_1, \ldots, x_n : s_n) : s \iff \ldots$ be some deterministic and case-complete algorithm which is well-formed for an admissible specification $S$, i.e. all function symbols (except $f$) used in the cases of $f$ are defined by $S$. Since $S$ is admissible it has a standard model $\mathcal{I}$. The standard model gives a meaning to each function symbol (except $f$) in the definition of $f$ because $f$ is well-formed for $S$. Here we are concerned with the termination of the algorithm $f$ and we have to define what this notion formally means.
If the interpreter has to evaluate a term \( f(q_{1,0}, \ldots, q_{n,0}) \) for some input \((q_{1,0}, \ldots, q_{n,0}) \in \mathcal{F(S)}_{q_{1,0}} \ldots_{q_{n,0}}\), some other term \( f(q_{1,1}, \ldots, q_{n,1}) \) stemming from a recursive call in \( f \) has to be evaluated, which in turn necessitates an evaluation of some further term \( f(q_{1,2}, q_{n,2}) \), etc. This means that the evaluation of each term \( t = f(q_{1,0}, \ldots, q_{n,0}) \) defines a so-called trace \((q_{1,0}, \ldots, q_{n,0}), (q_{1,1}, \ldots, q_{n,1}), (q_{1,2}, \ldots, q_{n,2}), \ldots\) of \( n \)-tuples of constructor ground terms which arise as evaluated arguments of \( f \) on the evaluation of \( t \). Intuitively, the algorithm \( f \) terminates if each trace is finite, i.e. during the evaluation of \( f(q_{1,0}, \ldots, q_{n,0}) \) some term \( f(q_{1,k}, q_{n,k}) \) is created which can be evaluated without a further recursion and therefore the evaluation of \( f(q_{1,0}, \ldots, q_{n,0}) \) must halt after finitely many steps.

The finiteness of a trace is formally captured by the notion of a well-founded relation: For a set \( M \), a binary relation \( <_M \) on \( M \) is well-founded iff there is no infinite sequence \( m_1, m_2, m_3, \ldots \) of elements in \( M \) such that \( \cdots <_M m_3 <_M m_2 <_M m_1 \). Now termination of \( f \) can be expressed by demanding the existence of some well-founded relation \( <_r \) on \( \mathcal{F(S)}_{q_{1,0}} \ldots_{q_{n,0}} \) such that \( (q_{1,i+1}, \ldots, q_{n,i+1}) <_r (q_{1,i}, \ldots, q_{n,i}) \) for each pair of adjacent terms in each trace which stems from the evaluation of any term \( f(q_{1,0}, \ldots, q_{n,0}) \).

The elements of a trace (except the first) arise from the evaluation of the arguments of some recursive calls in the algorithm \( f \). Since an algorithm may use nested recursions, i.e. terms of the form \( f(\ldots f(\ldots)) \), we cannot use the standard model \( \mathcal{I} \) of \( \mathcal{S} \) to evaluate the arguments of \( f(t_1, \ldots, t_n) \) if some argument \( t_i \) contains a further call of \( f \). Also the condition \( \varphi \) of a recursive case may contain recursive calls and then the truth of \( \varphi \) cannot be determined by \( \mathcal{I} \). Therefore the standard model is extended to an interpretation which can also be applied to \( f \)-terms:

Let \( \mathcal{I}_f \) be any standard interpretation for \( \Sigma \cup \{ f \} \) which coincides with \( \mathcal{I} \) for all function symbols in \( \Sigma \), i.e. \( \mathcal{I}_f \models \Phi(\Sigma) \), where \( \Sigma \) is the set of function symbols in \( \mathcal{S} \) and \( \Phi(\Sigma) \) are the axioms of \( \mathcal{S} \). This means that the universe of \( \mathcal{I}_f \) is exactly the set of constructor ground terms because "\( \mathcal{I}_f \) is standard" is demanded, \( \mathcal{I}_f \) "behaves" like \( \mathcal{I} \) for each function symbol in \( \mathcal{S} \) because \( \mathcal{I}_f \models \Phi(\Sigma) \) is demanded, and \( \mathcal{I}_f \) maps the function symbol \( f \) to some function defined on constructor ground terms, where arities etc. are respected of course. Each such interpretation is called an \( f \)-expansion of \( \mathcal{I} \), and termination means that (i) the standard model \( \mathcal{I}_f \) can be extended to an \( f \)-expansion \( \mathcal{I}_f = \Phi(\Sigma) \) that satisfies the definition of \( f \), i.e. \( \mathcal{I}_f \models \mathrm{DEF}_f \), and (ii) each trace which arises from an evaluation of a term \( f(\ldots) \) by \( \mathcal{I}_f \) is finite:

**Definition 3.1.** A case-complete and deterministic algorithm function \( f(x_1 : s_1, \ldots, x_n : s_n) : s \leftarrow \cdots \) terminates in an admissible specification \( S \) with standard model \( \mathcal{I} \) iff \( f \) is well-formed for \( S \) and there exists an \( f \)-expansion \( \mathcal{I}_f \) of \( \mathcal{I} \) and some well-founded relation \( <_r \) on \( \mathcal{F(S)}_{q_{1,0}} \ldots_{q_{n,0}} \) such that \( \mathcal{I}_f \models \mathrm{DEF}_f \) and for each \((q_1, \ldots, q_n) \in \mathcal{F(S)}_{q_{1,0}} \ldots_{q_{n,0}}\), for each recursive case "if \( \varphi \) then \( r \)" of \( f \) and for each term \( t_1, \ldots, t_n \) in \( \varphi \) or in \( r \)

\[
\mathcal{I}_f[x_1, \ldots, x_n/q_1, \ldots, q_n] \models \varphi \implies \\
\mathcal{I}_f[x_1, \ldots, x_n/q_1, \ldots, q_n](t_1, \ldots, t_n) <_r (q_1, \ldots, q_n).
\] (3.1)
Requirement (3.1) demands that for each \( n \)-tuple \((q_1 \ldots q_n)\) of constructor ground terms which satisfies the condition \( \varphi \) of a recursive case, i.e. \( \mathcal{F}_f[ x_1 \ldots x_n / q_1 \ldots q_n] \models \varphi \), the arguments \((t_1 \ldots t_n)\) of a recursive call \( f(t_1 \ldots t_n) \) are evaluated to some \( n \)-tuple \( \mathcal{F}_f[ x_1 \ldots x_n / q_1 \ldots q_n][t_1 \ldots t_n] \) of constructor ground terms which is smaller than the \( n \)-tuple \((q_1 \ldots q_n)\) with respect to the well-founded relation \( \prec \). So \( f \) decreases each input \((q_1 \ldots q_n)\) in each recursive call with respect to a well-founded relation, and by the well-foundedness of \( \prec \) each trace is finite and each evaluation must halt after finitely many steps. The following theorem is easily proved by noetherian induction upon \( \prec \):

**Theorem 3.2.** If a case-complete and deterministic algorithm function \( f(x_1 : s_1 \ldots x_n : s_n) : s \sqsubseteq \cdots \) is well-formed for an admissible specification \( S \) with standard model \( \mathcal{S} \), then \( f \) terminates in \( S \) iff there exists exactly one \( f \)-expansion \( \mathcal{F}_f \) of \( S \) such that \( \mathcal{F}_f \models \text{DEF}_f \).

So termination means that there is exactly one total function which satisfies the definition formulas of \( f \). We illustrate Definition 3.1 by some examples. Suppose that \( S \) contains the data structure \texttt{number} and consider the algorithm

\[
\text{function } f1(n : \texttt{number}) : \texttt{number} \equiv \\
\text{if } n = 0 \text{ then } 1 \\
\text{if } n \neq 0 \text{ then } f1(f1(n - 1)).^4
\]

We find that \( \mathcal{F}_{f1}(f1(0)) = 1 \) for each \( f1 \)-expansion which satisfies \( \text{DEF}_{f1} \). Hence requirement (3.1) demands in particular \( \mathcal{F}_{f1}(f1(0)) = 1 \prec 1 \) and \( \prec \) cannot be well-founded because there is no well-founded relation \( \prec \) such that \( m \prec m \) for some \( m \in M \). Consequently \( f1 \) does not terminate in \( S \). The algorithm

\[
\text{function } f2(n : \texttt{number}) : \texttt{number} \equiv \\
\text{if } n = 0 \text{ then } 0 \\
\text{if } n \neq 0 \text{ then } 2 + f2(f2(n - 1))
\]

does not terminate in \( S \) either: If \( \mathcal{F}_{f2} \) is an \( f2 \)-expansion such that \( \mathcal{F}_{f2}(f2(0)) \models \text{DEF}_{f2} \), then in particular \( \mathcal{F}_{f2} \models [f2(2) = 2 + f2(2)] \). But \( q \neq 2 + q \) for all numbers \( q \), i.e. there is no standard interpretation which satisfies \( \text{DEF}_{f2} \) and consequently the algorithm \( f2 \) does not terminate in \( S \).^5 But the algorithm

\[
\text{function } f3(n : \texttt{number}) : \texttt{number} \equiv \\
\text{if } n = 0 \text{ then } 0 \\
\text{if } n \neq 0 \text{ then } f3(f3(n - 1))
\]

---

4 Here we write e.g. \( 1 \) instead of \( \text{succ}(0) \), \( 1 + \cdots \) instead of \( \text{succ}(\ldots) \), \( -1 \) instead of \( \text{pred}(\ldots) \) etc. to ease readability.

5 Note that the non-termination of both algorithms has different reasons: more than one total function satisfies the definition of the algorithm \( f1 \) whereas no total function satisfies the definition formulas of the algorithm \( f2 \).
terminates in $S$ because the $f_3$-expansion $\mathcal{S}_{f_3}$ which assigns $f_3$ the constant mapping $0$ satisfies $\text{DEF}_{f_3}$ and with $\mathcal{S}_{f_3}(q) = q \lessdot 1 + q$ as well as $\mathcal{S}_{f_3}(f_3(q)) = 0 < 1 + q$ for each number $q$, where $\lessdot$ is the usual (well-founded) "less-than" relation on numbers, requirement (3.1) is satisfied.

Proving termination according to Definition 3.1 requires some intuition and skill because the required $f$-expansion $\mathcal{S}_f$ has to be found such that requirement (3.1) can be verified. Since $\mathcal{S}_f \models \text{DEF}_f$ is demanded, proving termination necessitates that the operation which is computed by the algorithm $f$ (or relevant properties of this operation at least) must be known to verify the termination of $f$. This means that generally reasoning about an algorithm's semantics is required for proving termination and correctness and termination have to be shown simultaneously (see e.g. the termination proofs for McCarthy's $\text{91}$-function and for Ashcroft's algorithm for list reversal [11, 12], for Takeuchi's function [16] or for the algorithm $\text{norm2}$ in [17]).

Since reasoning about an algorithm's semantics can be very difficult (and in particular for a machine) it seems worthwhile to look for some stronger termination requirement. Consider the following definition:

**Definition 3.3.** A case-complete and deterministic algorithm function $f(x_1 : s_1 \ldots x_n : s_n) : s \lessdot \cdots $ strongly terminates in an admissible specification $S$ with standard model $\mathcal{S}$ iff $f$ is well-formed for $S$ and there is some well-founded relation $\lessdot$ on $\mathcal{S}(\Sigma^*)_{s_1 \ldots s_n}$ such that for each $(q_1 \ldots q_n) \in \mathcal{S}(\Sigma^*)_{s_1 \ldots s_n}$, for each recursive case "if $\varphi$ then $r$" of $f$, for each term $f(t_1 \ldots t_n)$ in $\varphi$ or in $r$, and for each $f$-expansion $\mathcal{S}_f$

$$\mathcal{S}_f[x_1 \ldots x_n/q_1 \ldots q_n] \models \varphi \implies \mathcal{S}_f[x_1 \ldots x_n/q_1 \ldots q_n](t_1 \ldots t_n) \lessdot (q_1 \ldots q_n).$$  

(3.2)

For instance, the algorithms $\text{plus}$ and $\text{lt}$ from Section 2 strongly terminate. The difference between termination and strong termination is that Definition 3.3 does not demand the satisfiability of $\text{DEF}_f$ by a standard interpretation, but instead requirement (3.2) must hold for all $f$-expansions of $\mathcal{S}_f$. Consequently strong termination is independent of an algorithm's semantics.

It can be shown that strong termination entails termination (cf. [22, Lemma B.3.3]). But strong termination is not necessary for termination, i.e. there are terminating algorithms like $f_3$ which do not strongly terminate. Consider the algorithms $f_1$, $f_2$ and $f_3$ above. Since the recursive calls in these algorithms coincide, strong termination of $f_3$ would entail strong termination of $f_1$ and also of $f_2$ which in turn entails the termination of $f_1$ and $f_2$, i.e. a contradiction. Since requirement (3.2) must hold for all $f$-expansions, it must in particular also hold for those which do not satisfy $\text{DEF}_f$ and then the termination test may fail. Here there is no well-founded relation $\lessdot$ such that (3.2) holds for the $f_3$-expansion

---

*This also holds if other formal frameworks (which may even use partial functions) are used for proving termination, cf. [3, 11, 12, 17].*
\( J_{f_3} \) which assigns \( f_3 \) the successor function because then \( J_{f_3}(f_3(q)) = 1 + q \) for each number \( q \).

Strong termination is easier to verify than general termination because the satisfiability of \( \text{DEF}_f \) is not demanded. Instead only the recursive calls in an algorithm are considered, which means in particular that the result of a non-recursive case and the context of a recursive call are irrelevant for strong termination. So strong termination is not destroyed if an algorithm is modified by altering the result in a non-recursive case or the context of a recursive call. This does not hold for termination in general. E.g. algorithm \( f_1 \) differs from \( f_3 \) only in the result of the non-recursive case and \( f_2 \) differs from \( f_3 \) only in the context of the recursive call.

Strong termination provides the formal base on which our termination proof procedure rests. But since strong termination is only sufficient for termination it seems worthwhile to consider which terminating algorithms fail to terminate strongly. The following lemma gives some insight into the situation:

**Lemma 3.4.** If a case-complete and deterministic algorithm function
\[
 f(x_1:s_1\ldots x_n:s_n):s \leftarrow \cdots \quad \text{without recursive calls in the conditions of the cases and}
\]
\[
 \text{without nested recursions terminates in an admissible specification \( S \), then \( f \) strongly}
\]
\[
 \text{terminates in \( S \).}
\]

**Proof.** Requirement (3.1) holds for some \( f \)-expansion of \( J \) because \( f \) terminates. Since neither \( \varphi \) nor the argument \( t_i \) of a recursive call contain an \( f \)-term, (3.1) also holds for each \( f \)-expansion of \( J \). Hence \( f \) strongly terminates in \( S \).

We call an algorithm \( f \) normal iff \( f \) has no recursive calls in the conditions of the cases and has no nested recursions, and Lemma 3.4 shows that for normal algorithms strong termination is also necessary for termination. In other words, we generally fail in proving termination at most for algorithms with nested recursions or with recursions in conditions if we only test for strong termination.

But this restriction seems not too strong, because non-normal algorithms—although of theoretical interest—seem not to be very relevant in practice.\(^7\)

In this paper only normal algorithms are considered because this eases the required formalism. Strictly speaking this provides a further restriction of the class of algorithms which are considered for termination, because there are also non-normal but strongly terminating algorithms, e.g.

\[
 \text{function } f_4(n, m: \text{number}): \text{number} \leftarrow
\]
\[
 \text{if } n = 0 \text{ then } 1 + m
\]
\[
 \text{if } n \neq 0 \text{ then } f_4(n-1, f_4(n-1, m)).
\]

\(^7\) The above algorithm \( f_3 \), McCarthy’s 91-function, Takeuchi’s function, Ashcroft’s algorithm and Paulson’s algorithm norm2 are examples of terminating non-normal algorithms which do not strongly terminate.
Other examples for such algorithms are the usual algorithm for Ackermann's function (cf. [11, 12]), the algorithm value in [2] or the algorithm normif in [17]. The restriction to normal algorithms is only for the sake of the presentation and not necessitated by our proposal for proving termination (cf. [22]).

Subsequently we develop a termination proof procedure which generates a so-called termination hypothesis $\tau$ for each recursive call in a normal algorithm $f$ such that $\tau \in Th(S)$ entails requirement (3.2) of Definition 3.3. Each termination hypothesis is given to an induction theorem prover for verification, and if successful strong termination of $f$ is proved.

4. The estimation calculus

4.1. The size order

A frequently used well-founded order for termination proofs compares the size of the data objects under consideration. For instance, stacks are compared by their depth, lists by their length, trees by their number of nodes, etc. We obtain an abstract notion of size by counting the number of occurrences of reflexive constructors in a constructor ground term (where substructures are ignored). Formally the size of a constructor ground term is given by the $s$-size measure $\#_s : \mathcal{T}(\Sigma^*) \to \mathbb{N}$ which is defined as

$$\#_s(q) = \begin{cases} 0, & \text{if } q \notin \mathcal{T}(\Sigma^*) , \\ 0, & \text{if } q = \text{ircons}(\ldots) , \\ 1 + \sum_{i=1}^n \#_s(q_i), & \text{if } q = \text{rcons}(q_1 \ldots q_n) , \end{cases}$$

where $\text{ircons}$ is some irreflexive and $\text{rcons}$ is some reflexive constructor of $s$. Now comparing the $s$-sizes of a pair of constructor ground terms of sort $s$ with the usual $<_\mathbb{N}$ relation on the natural numbers $\mathbb{N}$, we obtain the size order $<_*$ which is a well-founded relation for the constructor ground terms of the data structure $s$.

Since we only count occurrences of $s$-constructors, substructures of the data structure $s$ are ignored. Hence we do not count the occurrences of $\text{succ}$ if we compare elements of the data structure $\text{list}$, and therefore

$$\text{add}(\text{succ}(\text{succ}(0)) \text{ empty}) <_* \text{add}(0 \text{ add}(0 \text{ empty}))$$

because

$$\#_{\text{list}}(\text{add}(\text{succ}(\text{succ}(0)) \text{ empty})) = 1 <_{\mathbb{N}} 2 = \#_{\text{list}}(\text{add}(0 \text{ add}(0 \text{ empty}))) .$$

All terms built with irreflexive constructors are minimal elements of the size order because only reflexive constructors are considered. Therefore $\text{empty}$ and $0$ are $<_*$-minimal constructor ground terms. For the $S$-expressions of LISP [15], which are defined in our notation by
**structure** atom(index : number)

nil cons (car : sexpr cdr : sexpr): sexpr,

we obtain nil and all terms of the form atom(...), as \(_{=,\text{minimal elements in }}\mathcal{F}(\Sigma, \mathcal{V})\).

Since we intend to use the size order to formulate and to verify the termination of algorithms, we extend our many-sorted language \(\mathcal{F}(\Sigma, \mathcal{V})\) by two new binary predicate symbols \(<_{,}\) and \(\leq_{,}\). The semantics of these predicate symbols are defined by

\[
\mathcal{J}[x^*/q^*] \models t <_{,} r \iff \\
\#_s(\mathcal{J}[x^*/q^*](t)) <_{,} \#_s(\mathcal{J}[x^*/q^*](r)),
\]

(4.1)

\[
\mathcal{J}[x^*/q^*] \models t \leq_{,} r \iff \\
\#_s(\mathcal{J}[x^*/q^*](t)) \leq_{,} \#_s(\mathcal{J}[x^*/q^*](r)),
\]

(4.2)

where \(\mathcal{J}\) is the standard model of an admissible specification \(S\), \(t\) and \(r\) are terms from \(\mathcal{F}(\Sigma, \mathcal{V})\), and \(x^*\) are the variables in \(t\) and \(r\) which are replaced by the constructor ground terms in \(q^*\). Now assume that \(\text{mod}\) computes the remainder operation (cf. Section 7) and consider the following algorithm for the computation of the greatest common divisor:

**function** gcd(n, m : number): number \(
\iff \\
\text{if } n = m \text{ then } n \\
\text{if } n \neq m \land n = 0 \text{ then } m \\
\text{if } n \neq m \land m = 0 \text{ then } n \\
\text{if } n \neq m \land n \neq 0 \land m \neq 0 \text{ then } \text{gcd(mod}(n \text{ mod}(m \text{ mod}(m n)))) \)

A termination requirement for \(\text{gcd}\) can be formulated in the extended many-sorted language as

\[
[\forall n, m : \text{number} \mod(n m) \leq_{,} n] \in \text{Th}(S),
\]

(4.3)

\[
[\forall n, m : \text{number} \mod(m n) \leq_{,} m] \in \text{Th}(S),
\]

(4.4)

\[
[\forall n, m : \text{number} n \neq m \land n \neq 0 \land m \neq 0 \\
\rightarrow \mod(n m) <_{,} n \lor \mod(m n) <_{,} m] \in \text{Th}(S)
\]

(4.5)

and our general problem is to verify statements of the form

\[
[\forall x^*: w \ t \leq_{,} r] \in \text{Th}(S),
\]

(4.6)

\[
[\forall x^*: w \varphi \rightarrow \cdots \lor t <_{,} r \lor \cdots] \in \text{Th}(S).
\]

(4.7)

A straightforward idea for proving those statements is to invent some set of axioms describing \(<_{,}\) as well as \(\leq_{,}\) and then to call an induction theorem prover for verification of (4.6) and (4.7). However, we do not follow this idea. Instead a calculus is developed, called the estimation calculus or the \(E\)-calculus for short,
which eases to prove those statements. The formulas of the E-calculus are called estimation formulas, where we define:

**Definition 4.1.** An expression of the form \( t \leq_{\Delta} r \) is an estimation formula iff \( t \) and \( r \) are terms from \( \mathcal{T}(\Sigma, \emptyset) \), and \( \Delta \) is a quantifier-free formula.

An estimation formula \( \langle t \leq_{\Delta} r \rangle \) is true, abbreviated \( \text{TRUE} \langle t \leq_{\Delta} r \rangle \), iff \( [\forall x^* : w t \leq_{r} r] \in \mathcal{Th}(S) \) and \( [\forall x^* : w \Delta \iff t \leq_{\Delta} r] \in \mathcal{Th}(S) \), where \( x^* \) are the variables in \( t, r \) and \( \Delta \).

We let \( \overline{\vdash} \langle t \leq_{\Delta} r \rangle \) denote that \( \langle t \leq_{\Delta} r \rangle \) can be derived in the E-calculus and we demand that the E-calculus is sound, i.e. \( \overline{\vdash} \langle t \leq_{\Delta} r \rangle \) implies \( \text{TRUE} \langle t \leq_{\Delta} r \rangle \). Then statements like (4.6) and (4.7) can be proved by showing

\[
\overline{\vdash} \langle t \leq_{\Delta} r \rangle, \quad \Phi(S) \cup I(S) \vdash [\forall x^* : w \varphi \rightarrow \cdots \vee \Delta \vee \cdots],
\]

using an induction theorem prover for (4.9). Since the E-calculus is sound, requirement (4.8) implies (4.6) as well as \( [\forall x^* : w \Delta \iff t \leq_{\Delta} r] \in \mathcal{Th}(S) \) and with requirement (4.9) statement (4.7) then also holds.

Of course, our approach seems somewhat indirect as compared to the method of using axioms to prove (4.6) and (4.7) directly. But as will become obvious later on, our method has certain advantages because \( \overline{\vdash} \) is decidable and trivial to compute and verification problems like (4.9) are usually much easier solved than verification problems like (4.7).

**4.2. Argument-bounded functions**

Suppose that \( [\forall x^* : w t \leq_{\Delta} r] \in \mathcal{Th}(S) \) has to be verified for a pair of terms \( t \) and \( r \) with variables \( x^* \). A common verification technique is to use estimations for proving those inequalities. Assume, for instance, we have to verify

\[
\forall x, y, z : \text{number} \quad (x - y)/z \leq_{\Delta} x \quad (4.10)
\]

where \( \leq_{\Delta} \) coincides with \( \leq_{\mathbb{N}} \) here. Assume further that we know certain estimation laws about the operations, viz. \( n/m \leq_{\Delta} n \) for the truncated quotient and \( (n - m) \leq_{\Delta} n \) for subtraction.\(^8\) Then we may conclude with the first law that \( (x - y)/z \leq_{\Delta} (x - y) \), with the second law we find that \( (x - y) \leq_{\Delta} x \), and with the transitivity of \( \leq_{\Delta} \) inequality (4.10) is proved. All we used in the proof were both estimation laws and the transitivity of \( \leq_{\Delta} \).

We formalize this technique for proving inequalities by defining a relation on terms which mirrors the semantical \( \leq_{\Delta} \) relation on the syntactical level. Therefore let us look closer at the estimation laws which were used in the example: Obviously both are very similar by stating that the size of a function's result is

\[^8\] We assume for the sake of the example that \( n/0 = 0 \) and \( 0 - m = 0 \).
always less than or equal to the size of its first argument. We call such functions 1-bounded. Of course, we may use other estimation laws stating that the result of a function is bounded by one of its arguments other than the first. We formalize this property with the notion of an argument-bounded function:

**Definition 4.2.** A function symbol \( f : s_1 \times \cdots \times s_n \rightarrow s \) is \( p \)-bounded iff \( 1 \leq p \leq n \) and

\[
[\forall x_1 : s_1 \ldots x_n : s_n \; f(x_1 \ldots x_n) \leq \_ x_p] \in \text{Th}(S).
\]

A function symbol \( f \) is argument-bounded iff \( f \) is \( p \)-bounded for some argument position \( p \) of \( f \).

For each \( p \in \mathbb{N}, \Gamma_p(S) \) denotes a set of \( p \)-bounded function symbols in an admissible specification \( S \) and \( \Gamma(S) \) is the family of all these sets. We write \( \Gamma_p \) and \( \Gamma \) if \( S \) is known from the context.

Argument-bounded functions are frequently used in computer science to define algorithms recursively. For instance, all reflexive selectors are 1-bounded provided they return an appropriate result if applied to a constructor to whom they do not belong (cf. Section 2.3). Hence \( \text{car}, \text{cdr}, \text{pred} \) and \( \text{tail} \) are 1-bounded if we assume e.g. \( \text{car}() = \text{cdr}() = \text{car}(\text{atom}(\ldots)) = \text{cdr}(\text{atom}(\ldots)) = \text{nil}, \; \text{tail}(\text{empty}) = \text{empty} \) and \( \text{pred}(0) = 0 \) because then \( \{\forall x : \text{number} \; \text{pred}(x) \leq \_ x, \forall x : \text{list} \; \text{tail}(x) \leq \_ x, \forall x : \text{sexpr} \; \text{car}(x) \leq \_ x, \forall x : \text{sexpr} \; \text{cdr}(x) \leq \_ x\} \subset \text{Th}(S) \).

If we define, for instance,

**function** \( \text{minus}(x, y : \text{number}) : \text{number} \leftarrow \)

\[
\text{if } x = 0 \lor y = 0 \text{ then } x
\]

\[
\text{if } x \neq 0 \land y \neq 0 \text{ then } \text{minus}(\text{pred}(x) \; \text{pred}(y))
\]

and

**function** \( \text{half}(x : \text{number}) : \text{number} \leftarrow \)

\[
\text{if } \text{pred}(x) = 0 \text{ then } 0
\]

\[
\text{if } \text{pred}(x) \neq 0 \text{ then } \text{succ}(\text{half}(\text{pred}(\text{pred}(x))))
\],

then \( \text{minus} \) and \( \text{half} \) are 1-bounded because \( \{\forall x, y : \text{number} \; \text{minus}(x, y) \leq \_ x\} \in \text{Th}(S) \) and \( \{\forall x : \text{number} \; \text{half}(x) \leq \_ x\} \in \text{Th}(S) \). Also the (truncated) quotient and the remainder operation are 1-bounded if we use definitions which guarantee \( \{\forall x : \text{number} \; \text{quotient}(x, 0) \leq \_ x\} \in \text{Th}(S) \) and \( \{\forall x : \text{number} \; \text{mod}(x, 0) \leq \_ x\} \in \text{Th}(S) \). Also the remainder is also 2-bounded if \( \{\forall x : \text{number} \; \text{mod}(x, 0) \leq \_ 0\} \in \text{Th}(S) \) is satisfied.\(^9\)

In COMMONLISP [20] for instance \( \text{nthcdr}, \text{member}, \text{intersection} \) and \( \text{remove} \)

\(^9\)We need such additional requirements only for exceptional arguments, e.g. \( \text{car}(\text{nil}) \) or division by zero, which would normally yield an undefined result. We would not need these requirements if we could use partial operations. But since the semantics of first-order logic demand total operations, we feel free for these cases to stipulate the values which are convenient for us. (Cf. Section 7.)
are 2-bounded, and intersection is also 1-bounded. Our algorithm remove from Section 1 also computes a 2-bounded function, and one can observe that all these functions are frequently used in recursive definitions of algorithms.

The notion of an argument-bounded function is the key concept for our formalization of estimation proofs: Given a family \( \Gamma \) of argument-bounded function symbols, a decidable estimation relation \( \leq_r \) on \( \mathcal{F}(\Sigma, \mathcal{V}) \) can be defined such that

\[
\forall x^* : w \ t_1 \leq_r t_n \in \text{Th}(S) \tag{4.11}
\]

for all terms \( t_1 \) and \( t_n \), where \( x^* \) is a list of all variables in \( t_1 \) and \( t_n \). The estimation relation \( \leq_r \) is based on the knowledge about data structures and the argument-bounded functions in \( \Gamma \). It provides a deductive requirement for the \( \leq_* \) relation, and the proof technique of estimation is mirrored by \( \leq_r \).

Suppose that \( [\forall x, y : \text{number} \ \text{minus}(\text{half}(\text{pred}(x))) \ \text{succ}(y)) \leq_* x] \in \text{Th}(S) \) has to be verified, where it is known that \( \text{minus} \), \( \text{half} \) and \( \text{pred} \) are 1-bounded, i.e. \( \{\text{minus}, \ \text{half}, \ \text{pred}\} \subset \Gamma \). Then we conclude that \( \text{minus}(\text{half}(\text{pred}(x))) \leq_r \text{half}(\text{pred}(x)) \) because \( \text{minus} \) is 1-bounded, we conclude that \( \text{half}(\text{pred}(x)) \leq_r \text{pred}(x) \) because \( \text{half} \) is 1-bounded, and we finally conclude that \( \text{pred}(x) \leq_r x \) because \( \text{pred} \) is 1-bounded. Therefore \( \text{minus}(\text{half}(\text{pred}(x))) \ \text{succ}(y)) \leq_r x \) is established as it was demanded. So the general idea to establish \( t_1 \leq_r t_n \) is to test whether \( t_n \) is a subterm of \( t_1 \), as \( x \) is a subterm of \( \text{minus}(\text{half}(\text{pred}(x))) \ \text{succ}(y)) \), where however only subterms in positions \( p \) are inspected for which a function is \( p \)-bounded.

### 4.3. Difference functions for argument-bounded functions

The estimation relation \( \leq_r \) only provides a deductive requirement for the (semantical) \( \leq_* \) relation, cf. requirement (4.11). But we need a deductive means for the strict \( <_* \) relation because this is the well-founded relation our termination proofs are based on. We therefore introduce the notion of a \( p \)-difference function:

**Definition 4.3.** A function symbol \( d : s_1 \times \cdots \times s_n \rightarrow \text{bool} \) is a \( p \)-difference function for a \( p \)-bounded function \( f : s_1 \times \cdots \times s_n \rightarrow s \) iff

\[
[\forall x_1 : s_1 \cdots x_n : s_n \ d(x_1 \cdots x_n) \leftrightarrow f(x_1 \cdots x_n) <_* x_p] \in \text{Th}(S) .
\]

A function \( d \) is a difference function for \( f \) iff \( d \) is a \( p \)-difference function for the \( p \)-bounded function \( f \).

Subsequently we assume the existence of a difference function for each argument-bounded function in \( \Gamma \) and we let \( \Delta^p f \) denote the \( p \)-difference function for \( f \in \Gamma_p \).

A function may be \( p \)-bounded for more than one argument position \( p \), as the function \( \text{min} : \text{number} \times \text{number} \rightarrow \text{number} \) for the minimum of a pair of numbers...
is 1- and also 2-bounded. For those functions a difference function is required for each such argument position \( p \), as \( \Delta^1 \text{min} \) and \( \Delta^2 \text{min} \) exist for \( \text{min} \).

If a \( p \)-bounded function \( f \) is applied to some input \( q_1 \ldots q_n \) such that it returns something \( <_* \)-smaller than \( q_p \), then the associated difference function \( \Delta^p f \) applied to the same input yields \( \text{true} \), and it returns \( \text{false} \) if \( f \) returns something of the same size as the input on argument position \( p \). For instance, we may define

\[
\text{function } \Delta^1 \text{minus}(x, y : \text{number}) : \text{bool} <\!
\begin{align*}
&\text{if } x = 0 \lor y = 0 \text{ then } \text{false} \\
&\text{if } x \neq 0 \land y \neq 0 \text{ then } \text{true}
\end{align*}
\]

\[
\text{function } \Delta^1 \text{pred}(x : \text{number}) : \text{bool} <\!
\begin{align*}
&\text{if } x = 0 \text{ then } \text{false} \\
&\text{if } x \neq 0 \text{ then } \text{true}
\end{align*}
\]

\[
\text{function } \Delta^1 \text{half}(x : \text{number}) : \text{bool} <\!
\begin{align*}
&\text{if } x = 0 \text{ then } \text{false} \\
&\text{if } x \neq 0 \text{ then } \text{true}
\end{align*}
\]

because then

\[
[\forall x, y : \text{number} \ \Delta^1 \text{minus}(x, y) \leftrightarrow \text{minus}(x, y) <_* x] \in \text{Th}(S),
\]

\[
[\forall x : \text{number} \ \Delta^1 \text{pred}(x) \leftrightarrow \text{pred}(x) <_* x] \in \text{Th}(S),
\]

\[
[\forall x : \text{number} \ \Delta^1 \text{half}(x) \leftrightarrow \text{half}(x) <_* x] \in \text{Th}(S).
\]

Using difference functions, a syntactical requirement for the strict \( <_* \) relation can be formulated. Consider again the estimation from Section 4.2

\[
\text{minus} (\text{half} (\text{pred}(x)) \text{ succ}(y)) \leq_\text{r} \text{half} (\text{pred}(x)) \leq_\text{r} \text{pred}(x) \leq_\text{r} x. \quad (4.12)
\]

If we can establish that at least one of the three inequalities is strict, i.e. (i) \( \text{minus} (\text{half} (\text{pred}(x)) \text{ succ}(y)) <_* \text{half} (\text{pred}(x)) \) or (ii) \( \text{half} (\text{pred}(x)) <_* \text{pred}(x) \) or (iii) \( \text{pred}(x) <_* x \), then (iv) \( \text{minus} (\text{half} (\text{pred}(x)) \text{ succ}(y)) <_* x \) is obviously proved.

The strictness of each estimation step can be expressed by a literal which is built with a difference function. We find e.g. \( \Delta^1 \text{minus} (\text{half} (\text{pred}(x)) \text{ succ}(y)) \) as an equivalent requirement for (i), we find \( \Delta^1 \text{half} (\text{pred}(x)) \) as an equivalent requirement for (ii) and we find \( \Delta^1 \text{pred}(x) \) as an equivalent requirement for the truth of (iii). Consequently

\[
\Delta^1 \text{minus} (\text{half} (\text{pred}(x)) \text{ succ}(y)) \lor \Delta^1 \text{half} (\text{pred}(x)) \lor \Delta^1 \text{pred}(x) \quad (4.13)
\]

is an equivalent requirement for the truth of (iv). Restated in a more readable

\[\text{We may define } \Delta^1 \text{min} (x, y) = \text{lt}(y, x) \text{ and } \Delta^2 \text{min} (x, y) = \text{lt}(x, y).\]
notation, it is inferred that \((x - 1)/2 - (y + 1) < x \iff (x - 1)/2 \neq 0 \land y + 1 \neq 0 \lor x - 1 \neq 0 \lor x \neq 0\) for all numbers \(x\) and \(y\).

So the general idea here is to scan an estimation \(t_1 \leq_r t_2 \leq_r \cdots \leq_r t_{n-1} \leq_r t_n\) step by step. For each estimation step \(t_i \leq_r t_{i+1}\), where \(t_i = f_i(t_i, t_{i+1}, \ldots)\) and \(t_{i+1}\) stands in position \(p_i\) of the \(p_i\)-bounded function \(f_i\), the corresponding “call” of the \(p_i\)-difference function \(\Delta^p_i f_i(t_i, t_{i+1}, \ldots)\) is collected. From all “calls” in the collection the disjunction

\[
\Delta^p_1 f_1(t_1, t_2, \ldots) \lor \Delta^p_2 f_2(t_2, t_3, \ldots) \lor \cdots \lor \Delta^p_{n-1} f_{n-1}(t_{n-1}, t_n),
\]
called the difference equivalent \(\Delta_r(t_1, t_n)\) of \(t_1\) and \(t_n\), is computed. Since each literal of \(\Delta_r(t_1, t_n)\) is built with a difference function, an equivalent requirement for \(t_1 <_r t_n\) is generated, and we have

\[
t_1 \leq_r t_n \implies [\forall x^* : w \Delta_r(t_1, t_n) \iff t_1 <_r t_n] \in \Theta(S)
\]
where \(x^*\) are the variables in \(t_1\) and \(t_n\).

4.4. Estimation rules based on argument-bounded functions

Knowing a family \(\Gamma\) of argument-bounded functions and the difference functions for them, our idea of proving inequalities by estimations can be formalized further: We define the axioms and the inference rules (both also called estimation rules) of the E-calculus such that inequalities (represented as estimation formulas) can be formally derived. The estimation rules combine the test for the estimation relation \(\leq_r\) and the computation of the difference equivalent \(\Delta_r\) such that

\[
t \leq_r r \iff |-_r \langle t \leq_r r, \Delta_r(t, r) \rangle
\]
holds. The E-calculus is “parameterized” with \(\Gamma\), i.e. \(|-_r \langle t \leq_r r, \Delta \rangle\) denotes that we derive an estimation formula by using only the information about the argument-boundedness of the functions given by \(\Gamma\).

On defining the estimation rules we have to care about the soundness of the calculus, i.e. that \(|-_r \langle t \leq_r r, \Delta \rangle\) implies \(\text{true} \langle t \leq_r r, \Delta \rangle\) (cf. Definition 4.1). Since \([\forall x^* : w \ t \leq_r t] \in \Theta(S)\) and \([\forall x^* : w \ \text{false} \iff t \leq_r t] \in \Theta(S)\) obviously holds for all terms \(t \in \mathcal{F}(\Sigma, \forall)\), \(\langle t \leq_r t, \text{false} \rangle\) is a true estimation formula and we may use

\[
\frac{-}{\langle t \leq_r t, \text{false} \rangle}
\]
for all \(t\)
as an axiom of the calculus, called the identity rule.

Now suppose that \(f : s_1 \times \cdots \times s_n \to s\) is a \(p\)-bounded function with \(p\)-difference function \(\Delta^p f : s_1 \times \cdots \times s_n \to \text{bool}\). Then for all terms \(t_1, \ldots, t_n\)

\[
\text{true} \langle f(t_1, \ldots, t_n) \leq_r t_p, \Delta^p f(t_1, \ldots, t_n) \rangle
\]
(4.16)
by Definitions 4.1, 4.2 and 4.3. Assume further that $\vdash_F \langle t_p \leq_s r, \Delta \rangle$. Then by the soundness of the E-calculus
\[
\text{TRUE} \langle t_p \leq_s r, \Delta \rangle.
\] (4.17)

From (4.16) and (4.17) we conclude
\[
\text{TRUE} \langle f(t_1, \ldots, t_n) \leq_s r, \Delta^f f(t_1, \ldots, t_n) \lor \Delta \rangle.
\] (4.18)

We formalize this reasoning step by an inference rule, called the argument estimation rule:
\[
\frac{\langle t_p \leq_s r, \Delta \rangle}{\langle f(t_1, \ldots, t_n) \leq_s r, \Delta^f f(t_1, \ldots, t_n) \lor \Delta \rangle} \quad \text{for all } f \in \Gamma_p(S), t_1, \ldots, t_n, r, \Delta.
\]

For instance, the 1-bounded function \textit{minus} with 1-difference function \Delta^1\textit{minus} yields the argument estimation rule
\[
\frac{\langle t_1 \leq_s r, \Delta \rangle}{\langle \text{minus}(t_1, t_2) \leq_s r, \Delta^1\text{minus}(t_1, t_2) \lor \Delta \rangle} \quad \text{for all } t_1, t_2, r, \Delta.
\]

With similar rules for \textit{half} and \textit{pred} the formal derivation in the E-calculus
\[
\langle x \leq_s x, \Delta^\text{false} \rangle
\]
\[
\langle \text{pred}(x) \leq_s x, \Delta^\text{pred}(x) \rangle^{11}
\]
\[
\langle \text{half}(\text{pred}(x)) \leq_s x, \Delta^\text{half}(\text{pred}(x)) \lor \Delta^\text{pred}(x) \rangle
\]
\[
\langle \text{minus}(\text{half}(\text{pred}(x)) \text{succ}(y)) \leq_s x, \Delta^\text{minus}(\text{half}(\text{pred}(x)) \text{succ}(y)) \lor \Delta^\text{half}(\text{pred}(x)) \rangle
\]
is obtained for our example from Sections 4.2 and 4.3, where we start with the identity rule and then use the argument estimation rule three times, viz. for \textit{pred}, \textit{half} and \textit{minus}.

4.5. Estimation rules based on data structures

We define some additional estimation rules to increase the performance of the E-calculus in proving inequalities. However, we will not obtain a complete calculus because we do not provide rules such that e.g. $\vdash_F \langle x \leq_s \text{plus}(x, y), y \neq 0 \rangle$. But we do not bother with this incompleteness because we use the calculus only for the special estimation problems necessitated by termination proofs and the computation of difference functions (cf. Sections 5 and 6). As will be demonstrated subsequently, the E-calculus is "complete enough" for our purposes.

All estimation rules introduced in this section depend only on the definition of

\[11\text{ We sometimes simplify } \varphi \lor \text{false} \to \varphi \text{ and } \varphi \lor \text{true} \to \text{true in our examples.}\]
the data structures, i.e. they can be uniformly obtained from the data structures in
a specification. Hence no specific knowledge is involved with these rules as
opposed to the argument estimation rule which presupposes that p-bounded
functions and also the p-difference functions for them are known.

We start with the axioms of the calculus, i.e. the estimation rules with an empty
list of premises: Consider a pair of terms \texttt{ircons}_1(...) and \texttt{ircons}_2(...) with
irreflexive constructors as leading function symbols, i.e. the s-size of both terms is
0. Then \([\forall \ldots \text{ircons}_1(...) \leq\texttt{ircons}_2(...)]) \in \text{Th}(S)\) and \([\forall \ldots \text{FALSE} \leftrightarrow
\text{ircons}_1(...) \prec\texttt{ircons}_2(...)]) \in \text{Th}(S)\) obviously holds, and it is sound to use

\[
\frac{}{\langle \text{ircons}_1(t_1 \ldots t_n) \leq\texttt{ircons}_2(r_1 \ldots r_m), \text{FALSE} \rangle}
\]

for all \(t_i, r_j\)

as an axiom, called the \textit{equivalence} rule. Now e.g. \(\vdash \langle \text{atom}(n) \leq\texttt{nil}, \text{FALSE} \rangle\)
holds.

Next we compare a term \texttt{ircons}(...) with a term \texttt{rcons}(...) where \texttt{ircons} is an
irreflexive constructor and \texttt{rcons} is a reflexive constructor, i.e. the s-size of
\texttt{ircons}(...) is 0 and the s-size of \texttt{rcons}(...) is 1 at least. Then \([\forall \ldots \text{ircons}(\ldots) \leq\texttt{rcons}(\ldots)]) \in \text{Th}(S)\) and \([\forall \ldots \text{TRUE} \leftrightarrow \text{ircons}(\ldots) \prec\texttt{rcons}(\ldots)]) \in \text{Th}(S)\) obviously
holds, and it is sound to use

\[
\frac{}{\langle \text{ircons}(t_1 \ldots t_n) \leq\texttt{rcons}(r_1 \ldots r_m), \text{TRUE} \rangle}
\]

for all \(t_i, r_j\)

as an axiom, called the \textit{strong estimation} rule. Now e.g. \(\vdash \langle \text{atom}(n) \leq\texttt{cons}(x\ y), \text{TRUE} \rangle\)
holds.

Now we compare a pair of terms \texttt{ircons}(...) and \texttt{r} of sort \(s\) such that \texttt{ircons} is an
irreflexive constructor, i.e. the s-size of \texttt{ircons}(...) is 0 and therefore \([\forall \ldots \text{ircons}(\ldots) \leq\texttt{r}]) \in \text{Th}(S)\). The s-size of \(r\) is 1 at least iff \(r = \texttt{rcons}_1(r)\ldots\texttt{rcons}_n(r)\) for some reflexive constructor \texttt{rcons}_j with selectors
\(\texttt{sel}_{j_1}(r) \ldots \texttt{sel}_{j_h}(r)\). Hence if \(\texttt{rcons}_1, \ldots, \texttt{rcons}_n\) are all reflexive constructors of \(s\),
then \([\forall \ldots (r = \texttt{rcons}_1(r) \ldots \texttt{sel}_{j_1}(r) \ldots \texttt{sel}_{j_h}(r)) \vee \ldots \vee r = \texttt{rcons}_n(r) \ldots \texttt{sel}_{j_1}(r) \ldots \texttt{sel}_{j_h}(r))] \leftrightarrow \text{ircons}(\ldots) \prec\texttt{rcons}(\ldots)\) \in \text{Th}(S)\), and it is sound to use

\[
\frac{}{\langle \text{ircons}(t_1 \ldots t_n) \leq\texttt{r}, r = \texttt{rcons}_1(\text{sel}_{j_1}(r) \ldots \texttt{sel}_{j_h}(r)) \ldots \texttt{rcons}_n(\text{sel}_{j_1}(r) \ldots \texttt{sel}_{j_h}(r)), \text{TRUE} \rangle}
\]

for all \(t_i, r\)

as an axiom, called the \textit{minimum} rule. Now e.g. \(\vdash \langle \text{atom}(n) \leq\texttt{cons}(x\ y),\ x =
\text{cons}(\text{car}(x)\ \text{cdr}(x)) \rangle\) holds.

We continue with the non-axiom rules: Suppose that some term \(t\) is compared
with \(\texttt{rcons}(r_1 \ldots r_k \ldots r_n)\), where \(k\) is a reflexive argument position of the
reflexive constructor \texttt{rcons} and therefore \([\forall \ldots r_k \prec\texttt{rcons}(r_1 \ldots t_k \ldots r_n)]] \in
\text{Th}(S)\). Suppose further that \(\vdash \langle t \leq\texttt{r}_k, \Delta \rangle\). Then \([\forall \ldots t \leq\texttt{r}_k ]] \in \text{Th}(S)\)
and therefore \([\forall \ldots \text{TRUE} \leftrightarrow t \prec\texttt{rcons}(r_1 \ldots r_k \ldots r_n)] \in \text{Th}(S)\), and it is sound to use
for all \( t, r_j, A \) as an inference rule, called the strong embedding rule. Now e.g. \( \vdash_r \langle \text{car}(y) \rangle \leq_r \text{cons}(x\ y) \), TRUE holds.

Finally we consider a pair of terms \( r\text{cons}(t_1 . . . t_n) \) and \( r\text{cons}(r_1 . . . r_n) \) where \( r\text{cons} \) is a reflexive constructor. If \( \vdash_r \langle t_j \leq_r r_j, A_j \rangle \) for all reflexive argument positions \( j \in \{ j_1, \ldots, j_h \} \) of \( r\text{cons} \), then \( \lbrack \forall \ldots r\text{cons}(t_1 . . . t_n) \leq_r r\text{cons}(r_1 . . . r_n) \rbrack \in \text{Th}(S) \) and \( \lbrack \forall \ldots (\Delta_{j_1} \lor \cdots \lor \Delta_{j_h}) \leftrightarrow r\text{cons}(t_1 . . . t_n) \leq_r r\text{cons}(r_1 . . . r_n) \rbrack \in \text{Th}(S) \) obviously holds. Therefore it is sound to use

\[
\langle t_1 \leq_r r_1, A_1 \rangle, \ldots, \langle t_h \leq_r r_h, A_h \rangle \leq_r r\text{cons}(t_1 . . . t_n) \leq_r r\text{cons}(r_1 . . . r_n), A_1 \lor \cdots \lor A_h
\]

as an inference rule, called the weak embedding rule. Now e.g. \( \vdash_r \langle \text{cons}(\text{car}(x)) \text{cdr}(y) \rangle \leq_r \text{cons}(x\ y), A_1 \text{car}(x) \lor A_1 \text{cdr}(y) \) holds.

A deduction of \( \langle t_n \leq_r r_n, A_n \rangle \) in the E-calculus, called an E-deduction for short, is a sequence of estimation formulas \( \langle t_1 \leq_r r_1, A_1 \rangle, \ldots, \langle t_n \leq_r r_n, A_n \rangle \) such that each estimation formula \( \langle t_k \leq_r r_k, A_k \rangle \) in the sequence either is an axiom or can be inferred by one of the inference rules from some estimation formulas preceding \( \langle t_k \leq_r r_k, A_k \rangle \) in the sequence.

\( \vdash_{\Gamma(S)} \langle t_n \leq_r r_n, A_n \rangle \) denotes the existence of such an E-deduction. The estimation relation \( t \leq_{\Gamma(S)} r \) abbreviates that \( \vdash_{\Gamma(S)} \langle t \leq_r r, A \rangle \) for some formula \( A \), and the difference equivalent \( \Delta_{\Gamma(S)}(t, r) \) of \( t \) and \( r \) is some formula \( A \) for which \( \vdash_{\Gamma(S)} \langle t \leq_r r, A \rangle \) holds.

It remains to verify the soundness of the estimation calculus:

**Theorem 4.4.** If \( \vdash_r \langle t \leq_r r, A \rangle \), then TRUE \( \langle t \leq_r r, A \rangle \).

**Proof.** Each axiom provides a true estimation formula and true estimation formulas are obtained from true estimation formulas by each inference rule as it was verified in Sections 4.4 and 4.5. Using these facts, the statement is easily proved by induction upon the length of the E-deduction. \( \square \)

### 4.6. A proof procedure for the estimation calculus

A proof procedure for the E-calculus is obtained by using the estimation rules in reverse direction. Each estimation rule can be written as

\[
\langle t_1 \leq_r r_1, A_1 \rangle, \ldots, \langle t_k \leq_r r_k, A_k \rangle \leq_r r, A \lor \Delta_1 \lor \cdots \lor \Delta_k
\]

\( \text{(ER)} \)

where FALSE must be substituted for the formula in \( \gamma \) in case of weak embedding to obtain the general form (ER) and the antecedent of the rule is empty in case of an axiom. From each estimation rule (ER), a production rule
is obtained and we have
\[ \vdash_r \langle t_1 \leq r, t_2, \Delta \rangle \text{ iff } (\langle t_1 \leq r, t_2, \text{FALSE} \rangle \Rightarrow^+ (\emptyset, \Delta)), \quad (4.19) \]
where \( \Rightarrow^+ \) is the reflexive and transitive closure of \( \Rightarrow \).

It can be easily observed from the definitions of the estimation rules that \( \{t_1 \leq r_1, \ldots, t_k \leq r_k\} \) has a strictly smaller number of variable and function symbols than \( \{t \leq r\} \) for each production rule (PR). Therefore \( \Rightarrow \) is \textit{noetherian}, i.e. there is no infinite derivation with respect to \( \Rightarrow \). Also only finitely many production rules can be applied to \( \{t \leq r\}, \Delta \), which means that \( \Rightarrow \) is \textit{locally finite}. Both properties of \( \Rightarrow \) imply that \( \vdash_r \langle t \leq r, \Delta \rangle \) is \textit{decidable}.

A proof procedure for the estimation calculus can be implemented by a production rule interpreter. This system requires some search control because \( \Rightarrow \) is not \textit{confluent}, i.e. there are terms \( t_1 \) and \( t_2 \) such that
\[ (E_1, A_1) \Rightarrow^* (E_2, A_2) \]
and
\[ (E_3, A_3) \Rightarrow^* (E_4, A_4) \]
for some \( (E_1, A_1) \) and \( (E_2, A_2) \), but
\[ (E_3, A_3) \Rightarrow^* (E_4, A_4) \]
for no \( (E_1, A_1) \), where \( \Rightarrow^* \) is the reflexive closure of \( \Rightarrow^+ \). So it may happen that
\[ (\emptyset, A_1) \Rightarrow^* (\emptyset, A_2) \]
as \( (\{0 \leq \text{succ}(x)\}, \Delta) \Rightarrow (\emptyset, \text{TRUE}) \) by strong estimation and \( (\{0 \leq \text{succ}(x)\}, \Delta) \Rightarrow (\emptyset, \text{succ}(x) = \text{succ}(\text{pred}(\text{succ}(x))) \text{ \text{v} } \Delta) \) by the minimum rule. This means that the computed difference equivalent depends upon the selection of the estimation rules. However this problem can be tolerated because \( A_1 \) and \( A_2 \) are equivalent by (4.19) and Theorem 4.4. A more serious problem arises if for all \( A' \)
\[ (\emptyset, A_1) \Rightarrow^* (\emptyset, A_2) \]
For instance, \( (\{\text{pred}(x) \leq \text{pred}(x)\}, \Delta) \Rightarrow (\emptyset, \Delta) \) by identity and \( (\{\text{pred}(x) \leq \text{pred}(x)\}, \Delta) \Rightarrow (x \leq \text{pred}(x), \Delta' \text{pred}(x) \text{ \text{v} } \Delta) \Rightarrow^* (\emptyset, A') \) if argument estimation is used. This shows that a dead end can be entered if the wrong estimation rule is selected.

We solve this problem by an appropriate interaction of the rules which minimizes backtracking (see [25] for details).

5. Generating termination hypotheses

5.1. Computing termination hypotheses with the E-calculus

Using the estimation relation \( \leq_r \) and the difference equivalent \( \Delta_r \), the generation of termination hypotheses is straightforward. If the algorithm tested for termination is given as \textit{function} \( f(x : s') : s \leftarrow \cdots \) and \( f(t) \) is a recursive call in a
case “if \( \varphi \) then \( r \)” of \( f \), then it is tested whether \( t \leq r \) \( x \) and if successful the termination hypothesis \([\forall x : s' \varphi \rightarrow \Delta_r (t, x)]\) is generated for this recursion. This is performed for each recursive call and each termination hypothesis is given to an induction theorem prover for verification. If a proof can be found, then \([\forall x : s' \varphi \rightarrow t <_r x] \in \text{Th}(S)\) holds for each recursion and consequently the algorithm \( f \) terminates.

E.g. for proving the termination of the algorithm \( \text{half} \) from Section 4.2, \( \vdash_r \langle \text{pred}(\text{pred}(x)) \leq_x x, \Delta^1 \text{pred}(\text{pred}(x)) \vee \Delta^1 \text{pred}(x) \rangle \) is deduced in the E-calculus and therefore the termination hypothesis

\[
\forall x : \text{number} \quad \text{pred}(x) \neq 0 \rightarrow \Delta^1 \text{pred}(\text{pred}(x)) \vee \Delta^1 \text{pred}(x) ,
\]  

(5.1)
i.e. \( x - 1 \neq 0 \rightarrow x - 1 \neq 0 \vee x \neq 0 \), is generated for \( \text{half} \). For the algorithm \( \text{sort} \) from the introduction \( \vdash_r \langle \text{remove}(\text{minimum}(x, x)) \leq_x x, \text{member}(\text{minimum}(x, x)) \rangle \) is computed, where \( \text{remove} \) is assumed as a 2-bounded function with 2-difference function \( \text{member} \). Hence the induction theorem prover is called with the termination hypothesis

\[
\forall x : \text{list} \quad x \neq \text{empty} \rightarrow \text{member}(\text{minimum}(x, x)) .
\]  

(5.2)

Our method is easily generalized for proving termination of algorithms with more than one argument: for an algorithm function \( f(x_1 : s_1 \ldots x_n : s_n) : s \leftarrow \ldots \) a set of termination positions \( P \subset \{1, \ldots, n\} \) is computed such that \( i \in P \) iff \( i \) is an argument position with \( t_i \leq r x_i \) for each recursive call \( f(t_1 \ldots t_n) \) in the algorithm. The termination test fails if \( P = \emptyset \). Otherwise a termination hypothesis \([\forall x_1 : s_1 \ldots x_n : s_n \varphi \rightarrow \bigvee_{i \in P} \Delta_r (t_i, x_i)\] is generated for each recursive call \( f(t_1 \ldots t_n) \) in a case “if \( \varphi \) then \( r \)” of \( f \). If each termination hypothesis can be proven, then in each recursive call at least one argument \( t_i \) is \( <_r \)-smaller than its corresponding initial input \( x_i \). By the definition of \( P \), no input \( x_i \) (with \( i \in P \)) is \( <_r \)-smaller than its corresponding argument \( t_i \) in any recursive call, and therefore the algorithm \( f \) must terminate.

Consider for instance the algorithm

\[
\text{function gcd2}(x, y : \text{number}) : \text{number} \leftarrow
\begin{align*}
& \text{if } x = 0 \vee y = 0 \text{ then } \text{max}(x, y) \\
& \text{if } x \neq 0 \land y \geq x \text{ then } \text{gcd2}(x \ \text{minus}(y, x)) \\
& \text{if } y \neq 0 \land y < x \text{ then } \text{gcd2}(\text{minus}(x, y), y) .
\end{align*}
\]

Since \( x \leq_r x \) as well as \( \text{minus}(y, x) \leq_r y \) in the first recursion and \( \text{minus}(x, y) \leq_r x \) as well as \( y \leq_r y \) in the second one, \( P = \{1, 2\} \) is computed (where \( \text{minus} \) is a 1-bounded function with difference function \( \Delta^1 \text{minus} \), cf. Section 4.3). Therefore

\[
\forall x, y : \text{number} \quad x \neq 0 \land y \geq x \rightarrow \text{false} \lor \Delta^1 \text{minus}(y, x) ,
\]  

(5.3)
i.e. \( x \neq 0 \land y \geq x \rightarrow y \neq 0 \land x \neq 0 \), is the computed termination hypothesis for the first recursion, and
\[ \forall x, y : \text{number} \quad y \neq 0 \land y < x \rightarrow \Delta^1 \text{minus}(x, y) \lor \text{false} , \]  
\[ (5.4) \]

i.e. \( y \neq 0 \land y < x \rightarrow x \neq 0 \land y \neq 0 \), is computed for the second one. We formalize this approach with the following termination criterion (see Appendix B for the proof):

**Theorem 5.1.** A normal algorithm function \( f(x_1 : s_1, \ldots, x_n : s_n) : s \leq \cdots \) strongly terminates in an admissible specification \( S \) if there exists some non-empty set \( P \subset \{1, \ldots, n\} \) such that for each recursive call \( f(t_1, \ldots, t_n) \) in a case "if \( \varphi \) then . . . " of \( f \):

1. \( t_i \leq_{f(S)} x_i \) for all \( i \in P \), and
2. \( [\forall x_1 : s_1, \ldots, x_n : s_n \; \varphi \rightarrow \bigvee_{i \in P} \Delta^1_{f(S)}(t_i, x_i)] \in \text{Th}(S) \).

The formulas of requirement (2) in Theorem 5.1 are called the termination hypotheses for \( f \). The termination hypotheses are the weakest requirements for the termination of an algorithm with respect to each syntactical termination criterion based on the size order, because with Theorem 4.4 the difference equivalent is sufficient and necessary for a strict decrease of an argument in a recursive call.\(^{12} \)

The set \( P \) in Theorem 5.1, called a measured subset in [2], denotes the argument positions which are relevant for the termination of an algorithm. If requirements (1) and (2) of Theorem 5.1 are satisfied for an algorithm \( f \), then \( \tau(x_1, \ldots, x_n) = \sum_{i \in P} \Delta^1_{f(S)}(x_i) \) is a termination function for \( f \) as it can be seen from the proof of Theorem 5.1.

We may weaken the requirements of Theorem 5.1 by comparing the argument list of a recursive call with the list of initial arguments by some lexicographical order based on the size order. This modification of the termination criterion is required for proving termination of algorithms like Ackermann’s function, cf. [22] for the generalized termination criterion.\(^{13} \) Also termination requirements based on other tuple-orders built with the size order, e.g. the multiset order [6], may be formulated if required.

For an implementation the set \( P \) in Theorem 5.1 should be chosen as the maximal set of indices which satisfies requirement (1) of Theorem 5.1 as the following example reveals. Consider the algorithm \( \text{gcd} \) from Section 4.1, where \( \text{mod} \) (cf. Section 7) is a 1-bounded function with 1-difference function

\[
\text{function} \; \Delta^1 \text{mod}(n, m : \text{number}) : \text{bool} \leftarrow \\
\text{if} \; n = 0 \land m = 0 \; \text{then} \; \text{false} \\
\text{if} \; n \neq 0 \land m = 0 \; \text{then} \; \text{true} \\
\text{if} \; n < m \; \text{then} \; \text{false}
\]

\(^{12} \) This corresponds to the requirement in [2], that a human user states each induction lemma with the "cleanest hypotheses available" in order to obtain a termination hypothesis as weak as possible. See also Section 8.

\(^{13} \) We found no frequent need for lexicographical comparisons. For instance, Ackermann’s function is the only algorithm in [2] which requires a lexicographical argumentation for proving termination.
if \( n \geq m \land m \neq 0 \) then true.

Since \( \text{mod}(n \ m) \leq_r n \) we may define \( P = \{1\} \), and then the false termination hypothesis

\[
\forall n, m : \text{number} \ n \neq m \land n \neq 0 \land m \neq 0 \rightarrow \Delta^1 \text{mod}(n \ m),
\]

i.e. \( n \neq m \land n \neq 0 \land m \neq 0 \rightarrow n \geq m \land m \neq 0 \), is obtained. Alternatively \( P = \{2\} \) can be used because \( \text{mod}(m \ n) \leq_r m \). However then the false termination hypothesis

\[
\forall n, m : \text{number} \ n \neq m \land n \neq 0 \land m \neq 0 \rightarrow \Delta^1 \text{mod}(m \ n),
\]

i.e. \( n \neq m \land n \neq 0 \land m \neq 0 \rightarrow m \geq n \land n \neq 0 \), is computed. But using \( P = \{1, 2\} \) the true termination hypothesis

\[
\forall n, m : \text{number} \ n \neq m \land n \neq 0 \land m \neq 0 \rightarrow \Delta^1 \text{mod}(n \ m) \lor \Delta^1 \text{mod}(m \ n),
\]

i.e. \( n \neq m \land n \neq 0 \land m \neq 0 \rightarrow n \geq m \land m \neq 0 \lor m \geq n \land n \neq 0 \), is obtained for \( \gcd \).

5.2. Improvements of the termination test

The central idea for proving termination is to deduce \( \Delta^r (t, x, z(t, x)) \), where \( x \) is some formal parameter of an algorithm and \( t \) is the corresponding actual parameter in a recursive call under some condition \( \varphi \), and then to prove a termination hypothesis of the form \( [\forall \cdots \varphi \rightarrow \cdots \lor \Delta^r (t, x) \lor \cdots] \). However it may happen that \( [\forall \cdots \varphi \rightarrow \cdots \lor t \leq_r x \lor \cdots] \in T h(S) \) but \( \not \forall^r (t \leq_r x, \dots) \). Sometimes the condition \( \varphi \) has to be considered not only for proving the termination hypothesis but also for generating it. The following algorithm (computing the minimal element of a non-empty list) provides an example:

```plaintext
function minimum(x : list) : number
if x = empty then 0
if x = add(head(x) tail(x)) \land tail(x) = empty then head(x)
if x = add(head(x) tail(x)) \land tail(x) = add(head(tail(x)) tail(tail(x)))
\land head(x) > head(tail(x)) then minimum(tail(x))
if x = add(head(x) tail(x)) \land tail(x) = add(head(tail(x)) tail(tail(x)))
\land head(x) < head(tail(x)) then
minimum(add(head(x) tail(tail(x)))) .
```

Obviously, \( \text{minimum} \) terminates but this cannot be verified by our method. This is because the estimation

\[
\text{add}(\text{head}(x) \ \text{tail}(\text{tail}(x))) \leq_r x
\]

does not hold and therefore the termination test fails for the second recursion of \( \text{minimum} \). But since
\[ x = \text{add}(\text{head}(x) \; \text{tail}(x)) \land \text{tail}(x) = \text{add}(\text{head}(\text{tail}(x)) \; \text{tail}(\text{tail}(x))) \]

is demanded for this case we may test for
\[ \text{add}(\text{head}(x) \; \text{tail}(\text{tail}(x))) \leq_{r} \text{add}(\text{head}(x) \; \text{add}(\text{head}(\text{tail}(x)) \; \text{tail}(\text{tail}(x)))) \quad (5.9) \]

instead. Now with \( \vdash_{r} (\text{add}(\text{head}(x) \; \text{tail}(\text{tail}(x))) \leq_{*} \text{add}(\text{head}(x) \; \text{add}(\text{head}(\text{tail}(x)) \; \text{tail}(\text{tail}(x)))) \), \text{TRUE} \) the trivial termination hypothesis \([\forall x: \text{list} \ldots \rightarrow \text{TRUE}] \) is generated for the second recursion.

The general problem is that we ignore the bindings imposed by the condition \( \varphi \) of a recursive case on the parameter \( x \) when we test for \( t \leq_{r} x \). It may happen that such a condition expresses the shape of a formal parameter in terms of constructors and this information is sometimes required to compute a termination hypothesis. Consequently the termination test fails if this information is ignored, as it did for the algorithm \textit{minimum}.

Our solution for this problem is to replace an actual parameter \( t \) and a formal parameter \( x \) by their so-called representatives \( [t]_{\varphi} \) and \( [x]_{\varphi} \), where
\[ [\forall \cdots \varphi \rightarrow [r]_{\varphi} = r] \in Th(S) \quad (5.10) \]

for each term \( r \) and its representative \( [r]_{\varphi} \). Now it is tested for \( [t]_{\varphi} \leq_{r} [x]_{\varphi} \) and if successful the termination hypothesis
\[ [\forall \cdots \varphi \rightarrow \cdots \lor \Delta([t]_{\varphi}, [x]_{\varphi}) \lor \cdots] \quad (5.11) \]

is computed. This approach is sound because a proof of (5.11) entails \([\forall \cdots \varphi \rightarrow \cdots \lor [t]_{\varphi} \leq_{*} [x]_{\varphi} \lor \cdots] \in Th(S) \) and (5.10) then implies \([\forall \cdots \varphi \rightarrow \cdots \lor t \leq_{*} x \lor \cdots] \in Th(S) \).

For computing representatives we demand that each condition of a case is given as a conjunction of literals and certain inconsistencies are forbidden to guarantee the uniqueness of each representative. For instance, 0 and \( \text{succ}(\text{pred}(x)) \) both could be computed as representatives of \( x \) if (inconsistent) conditions like \( x = 0 \land x = \text{succ}(\text{pred}(x)) \) would be allowed.

**Definition 5.2.** A case "if \( \varphi \; \text{then} \ldots \)" of an algorithm is in conjunctive normal form (\textit{cnf} for short) iff \( \varphi \) is a conjunction of literals.

A case "if \( \varphi \; \text{then} \ldots \)" in \textit{cnf} is structural consistent iff \( \varphi \) does not contain a pair of unnegated equations \( q = c_{1}(\ldots) \) and \( q = c_{2}(\ldots) \) such that \( c_{1} \) and \( c_{2} \) are constructor functions.

An algorithm \( f \) is in conjunctive normal form (\textit{cnf}) iff each case of \( f \) is in \textit{cnf}, and \( f \) is in structure normal form (\textit{snf}) iff \( f \) is in \textit{cnf} and each case of \( f \) is structural consistent.

Each algorithm can be transformed into an equivalent \textit{snf} algorithm (cf. [22]). As we have seen in the above example, the representative of \( x \) is obtained by applying equations of the condition of the algorithm's case: Using \( x = \text{add}(\text{head}(x) \)
tail(x) we first replace \( x \) with \( add(head(x) \ tail(x)) \), and then obtain \( add(head(x) \ add(head(tail(x)) \ tail(tail(x)))) \) by using \( tail(x) = add(head(tail(x)) \ tail(tail(x))) \). But we use only specific equations, called structure equations subsequently, which express the shape of a term in terms of constructors. We do not need arbitrary equations but only structure equations for our purposes, because only structure equations support the estimations necessary for the termination test.

**Definition 5.3.** An equation of the form \( t = cons(sel_1(t) \ldots sel_n(t)) \) is a structure equation iff \( cons \) is a constructor function and \( sel_1, \ldots, sel_n \) are the selectors of \( cons \).

The equations \( x = empty, x = add(head(x) \ tail(x)), tail(x) = add(head(tail(x)) \ tail(tail(x))), n = succ(pres(n)), pres(n) = succ(pres(pres(n))) \) and \( y = cons(car(y) \ cdr(y)) \) are examples for structure equations. Structure equations \( t = cons(sel_1(t) \ldots sel_n(t)) \) are used to compute the representative of a term \( r \) by replacing an occurrence of \( t \) in \( r \) by the right-hand side \( cons(sel_1(t) \ldots sel_n(t)) \) of the structure equation. Hence the right-hand side of a structure equation designates the term which replaces another term for the computation of a representative.

However, structure equations must be applied in a restricted way for obtaining useful results. For instance, we may replace \( x \) by \( add(head(x) \ tail(x)) \) using the structure equation \( x = add(head(x) \ tail(x)) \), then replace the new term by \( add(head(add(head(x) \ tail(x))) \ tail(add(head(x) \ tail(x)))) \) using the same structure equation again and so on. To avoid such replacements we also have to designate the terms which we want to replace. We call these candidates for a replacement parameter components, i.e. a parameter component is a term which may be replaced by another term when a representative is computed.

As a starting point we declare each formal parameter of an algorithm as a parameter component. Hence we may replace \( x \) with \( add(head(x) \ tail(x)) \) using the structure equation \( x = add(head(x) \ tail(x)) \). Now the shape of \( add(head(x) \ tail(x)) \) can be refined in terms of constructors if \( head(x) \) or \( tail(x) \) can be replaced by the right-hand side of a structure equation. Consequently \( head(x) \) and \( tail(x) \) are declared as parameter components, and since \( tail(x) = add(head(tail(x)) \ tail(tail(x))) \) is a structure equation in our example, we obtain \( add(head(x) \ add(head(tail(x)) \ tail(tail(x)))) \). We continue with \( head(tail(x)) \) and \( tail(tail(x)) \) as a parameter components. But since there are no structure equations of the form \( head(tail(x))) = \cdots \) and \( tail(tail(x)) = \cdots \), the computation stops with \( add(head(x) \ add(head(tail(x)) \ tail(tail(x)))) \) as a representative of \( x \). This technique is formalized by the following two definitions:

**Definition 5.4.** Let "if \( \varphi \) then..." be a case of an algorithm function \( f(x_1 : s_1 \ldots x_n : s_n) : s \leftarrow \cdots \) in snf. Then each formal parameter \( x_i \) is a parameter component of \( \varphi \), and if \( t \) is a parameter component of \( \varphi \) and \( \varphi \) contains an unnegated structure equation \( t = cons(sel_1(t) \ldots sel_n(t)) \), then \( sel_1(t), \ldots, sel_n(t) \) are also parameter components of \( \varphi \).
Definition 5.5. Let "if \( \varphi \) then..." be a case of an algorithm function
\[ f(x_1 : s_1 \ldots x_n : s_n) : s \leftarrow \cdots \text{ in snf}. \]
Then the representative of a term \( t \) with respect to \( \varphi \), abbreviated \([t]_{\varphi}\), is defined as
\[
[t]_{\varphi} = \begin{cases} 
\text{cons}([sel_1(t)]_{\varphi} \ldots [sel_n(t)]_{\varphi}), & \text{if } t = \text{cons}(sel_1(t) \ldots sel_n(t)) \text{ is a structure equation in } \varphi \text{ and } t \text{ is a parameter component of } \varphi, \\
g([t_1]_{\varphi} \ldots [t_n]_{\varphi}), & \text{if } t = g(t_1 \ldots t_n), \text{ where } g \text{ is any function symbol, and } t \text{ is not a parameter component of } \varphi, \\
t, & \text{otherwise.}
\end{cases}
\]

Since we define representatives only with algorithms in structure normal form, each case is structural consistent and therefore at most one structure equation \( t = \text{cons}(sel_1(t) \ldots sel_n(t)) \) exists for each parameter component \( t \). The requirements in Definition 5.5 also exclude each other. Consequently each representative \([t]_{\varphi}\) is uniquely determined. Subsequently we omit subscripts and write \([t]_{\varphi}\) instead of \([t]_{\varphi} \) if \( \varphi \) is obvious.

The computation of representatives can be viewed as a special equational reasoning procedure tailored for our purposes. Since we only use (structure) equations when computing the representative of a term, each term must denote the same object as its representative if the condition of the case is satisfied:

Lemma 5.6. If "if \( \varphi \) then..." is a case of an algorithm in snf, then \([\forall \cdots \varphi \rightarrow [r]_{\varphi} = r]\) \( \in \text{Th}(S) \) for all terms \( r \).

Using term representatives we now improve our termination test by the following corollary which is an obvious consequence of Theorem 5.1 and Lemma 5.6:

Corollary 5.7. A normal algorithm function \( f(x_1 : s_1 \ldots x_n : s_n) : s \leftarrow \cdots \text{ in snf} \) strongly terminates in an admissible specification \( S \) if there exists some non-empty set \( P \subset \{1, \ldots, n\} \) such that for each recursive call \( f(t_1, \ldots t_n) \) in a case "if \( \varphi \) then..." of \( f \)
\begin{enumerate}
\item \([t_i]_{\varphi} \leq f(S) [x_i]_{\varphi} \) for all \( i \in P \), and
\item \([\forall x_1 : s_1 \ldots x_n : s_n \varphi \rightarrow \bigvee_{i \in P} \Delta f(S)([t_i]_{\varphi}, [x_i]_{\varphi})] \in \text{Th}(S)\).
\end{enumerate}

Note that we must also use the representatives \([t_i]_{\varphi}\) of the actual parameters in a recursive call, because if only the representatives \([x_i]_{\varphi}\) of the formal parameters are used, termination cannot be verified for certain algorithms. Consider for instance the algorithm

```plaintext
function log(n : number) : number \leftarrow
if n = 0 then 0
if n = succ(pred(n)) \land pred(n) = 0 then 0
if n = succ(pred(n)) \land pred(n) = succ(pred(pred(n))) then
    succ(log(half(n))) .
```
Since $\frac{1}{2}(n) \not\equiv_r \text{succ}(\text{succ}(\text{pred}(\text{pred}(n)))) = [n]$, we cannot compute a termination hypothesis if we only use the representative of the formal parameter $n$. But if the representative is also computed for the actual parameter, then $[\frac{1}{2}(n)] = \frac{1}{2}([n])] = \cdots$ is obtained (cf. Definition 5.5) and with $\vdash_r (\frac{1}{2}([n])) <_r [n], \Delta^1(\frac{1}{2}([n]))$ the termination hypothesis

$$\forall n : \text{number} \quad n = \text{succ}(\text{pred}(n)) \land \text{pred}(n) = \text{succ}(\text{pred}(\text{pred}(n)))$$

$$\rightarrow \Delta^1(\frac{1}{2}(\text{succ}(\text{pred}(\text{pred}(n))))),$$

(5.12)
i.e. $n \neq 0 \land n - 1 \neq 0 \rightarrow 2 + (\ldots) \neq 0$, is computed for log.

5.3. Algorithms in positive structure normal form

The success of the termination criterion given by Corollary 5.7 depends on the form of the algorithm which is tested for termination. If we avoid structure equations in the condition of an algorithm, the method may fail as the following examples reveal: If the condition in the second recursive case of the algorithm minimum is given as

$$\text{tail}(x) = \text{add}(\text{head}(\text{tail}(x)) \text{tail}(\text{tail}(x))) \land \text{head}(x) \leq \text{head}(\text{tail}(x))$$

(5.13)
or is given as

$$\text{tail}(x) \not\equiv \text{empty} \land \text{head}(x) \leq \text{head}(\text{tail}(x))$$

(5.14)

then $x = [x]$ and the termination test fails. The reason for this failure is that the required structure equations are not given in the conditions, but are expressed implicitly instead.

For instance, $\text{tail}(x) = \text{add}(\text{head}(\text{tail}(x)) \text{tail}(\text{tail}(x)))$ entails $x = \text{add}(\text{head}(x) \text{tail}(x))$ and with this additional structure equation

$$\text{add}(\text{head}(x) \text{tail}(\text{tail}(x)))$$

$$\not\equiv_r \text{add}(\text{head}(x) \text{add}(\text{head}(\text{tail}(x)) \text{tail}(\text{tail}(x)))) = [x]$$

holds for (5.13). Also $\text{tail}(x) \not\equiv \text{empty}$ entails $x = \text{add}(\text{head}(x) \text{tail}(x))$ as well as $\text{tail}(x) = \text{add}(\text{head}(\text{tail}(x)) \text{tail}(\text{tail}(x)))$, and using these structure equations the required representative for $x$ can be computed also for (5.14). To avoid an implicit representation of structure equations a further normal form for algorithms is defined:

**Definition 5.8.** An algorithm function $f(x_1 : s_1 \ldots x_n : s_n) : s \subseteq \cdots$ is in *positive structure normal form* (psnf) iff $f$ is in snf and for each case "if $\varphi$ then..." in $f$:

1. $\varphi$ does not contain a subformula of the form $t \not\equiv \text{cons}(\ldots)$, and
2. each equation $t = \text{cons}(\ldots)$ in $\varphi$ is a structure equation such that
3. $t$ is a parameter component of $\varphi$,

where $\text{cons}$ is any constructor function.
For instance, the algorithms \texttt{minimum} and \texttt{log} are in psnf, but \texttt{half} is not a psnf algorithm.

By requirement (1), a case like "if \( x \neq \text{add}(y, z) \) then \( r \)" is not allowed and must be replaced by the sequence of cases

\begin{align*}
  &\text{if } x = \text{empty} \text{ then } r, \\
  &\text{if } x = \text{add}(\text{head}(x), \text{tail}(x)) \land \text{head}(x) \neq y \text{ then } r, \\
  &\text{if } x = \text{add}(\text{head}(x), \text{tail}(x)) \land \text{head}(x) = y \land \text{tail}(x) \neq z \text{ then } r
\end{align*}

to obtain an equivalent psnf algorithm.

A case like "if \( x = \text{add}(y, z) \) then \( r \)" is forbidden by requirement (2) and must be replaced by

\begin{align*}
  &\text{if } x = \text{add}(\text{head}(x), \text{tail}(x)) \land \text{head}(x) = y \land \text{tail}(x) = z \text{ then } r.
\end{align*}

Requirement (3) does not allow cases like

\begin{align*}
  &\text{if } \text{tail}(x) = \text{add}(\text{head}(\text{tail}(x)), \text{tail}(\text{tail}(x))) \text{ then } r,
\end{align*}

which are modified to

\begin{align*}
  &\text{if } x = \text{add}(\text{head}(x), \text{tail}(x)) \land \text{tail}(x) = \text{add}(\text{head}(\text{tail}(x)), \text{tail}(\text{tail}(x))) \text{ then } r
\end{align*}

for obtaining an equivalent psnf algorithm. Also cases like "if \( \text{tail}(x) = \text{empty} \) then \( r \)" are not allowed and must be replaced by a sequence of cases:

\begin{align*}
  &\text{if } x = \text{empty} \text{ then } r , \\
  &\text{if } x = \text{add}(\text{head}(x), \text{tail}(x)) \land \text{tail}(x) = \text{empty} \text{ then } r.
\end{align*}

These examples illustrate how snf algorithms are transformed into equivalent psnf algorithms for increasing the success of the termination test.

6. Argument-bounded algorithms

6.1. Introduction

The generation of termination hypotheses is based on the argument-bounded functions in \( \Gamma \) and on their difference functions and one may wonder where the essential information actually comes from?

Often an argument-bounded function \( g \) can be recognized and a difference function for \( g \) can be found by machine. The starting point is trivial: Since each reflexive selector \( rsel \) of a data structure \( s \) is 1-bounded (cf. Sections 2.3 and 4.2), \( rsel \) is inserted into \( \Gamma_1 \) and a so-called 1-difference algorithm \textbf{function} \( \Delta^1 rsel(x:s) : \text{bool} \leftarrow \cdots \) is synthesized, which contains a case

\begin{align*}
  &\text{if } x = \text{ircons}(sel_1(x) \ldots sel_n(x)) \text{ then false}
\end{align*}

for each irreflexive constructor \( \text{ircons} \) of \( s \), and contains a case
if \( x = \text{rcons}(\text{sel}_1(x) \ldots \text{sel}_n(x)) \) then true

for each reflexive constructor \( \text{rcons} \) of \( s \), where \( \text{sel}_1, \ldots, \text{sel}_n \) are the selectors of \( \text{ircons} \) or \( \text{rcons} \) respectively. We obtain for instance for the data structures \( \text{list} \) and \( \text{sexpr} \)

function \( \Delta^1\text{tail}(x : \text{list}) : \text{bool} \leftarrow \)

if \( x = \text{empty} \) then false
if \( x = \text{add}(\text{head}(x) \text{tail}(x)) \) then true

function \( \Delta^1\text{cdr}(x : \text{sexpr}) : \text{bool} \leftarrow \)

if \( x = \text{atom}(\text{index}(x)) \) then false
if \( x = \text{nil} \) then false
if \( x = \text{cons}(\text{car}(x) \text{cdr}(x)) \) then true

and \( \Delta^1\text{car} \) is defined like \( \Delta^1\text{cdr} \).

Since function \( \Delta^1\text{rsel}(x : s) : \text{bool} \leftarrow \cdots \) is deterministic, case-complete and (vacuously) terminates, each admissible specification remains admissible when extended by the 1-difference algorithm of a reflexive selector. So we we may assume that each admissible specification \( S \) contains the 1-difference algorithms for all reflexive selectors of the data structures in \( S \), and then \( \Delta^1\text{rsel} \) is a 1-difference function for \( \text{rsel} \).

To recognize other argument-bounded functions than reflexive selectors we define a decidable algorithm schema such that each instance of this schema (called a \( p \)-bounded algorithm) computes a \( p \)-bounded function. So for each algorithm function \( g(x^* : w) : s \leftarrow \cdots \) in an admissible specification it can be decided whether \( g \) is an instance of this algorithm schema. If the instance test succeeds, \( g \) must be a \( p \)-bounded function and is inserted into \( \Gamma_p \).

Then a recursive synthesis schema is used to synthesize a so-called \( p \)-difference algorithm function \( \Delta^p g(x^* : w) : \text{bool} \leftarrow \cdots \) for the algorithm just recognized as \( p \)-bounded. Since each \( p \)-difference algorithm is deterministic, case-complete and terminates "by construction", each admissible specification remains admissible when extended by a \( p \)-difference algorithm. So we may assume that each admissible specification \( S \) contains a \( p \)-difference algorithm for each \( p \)-bounded algorithm \( g \) in \( S \), and then \( \Delta^p g \) is a \( p \)-difference function for \( g \) as guaranteed by the synthesis procedure.

The main idea embodied in the algorithm schema is to construct a meta-induction proof for verification that the function computed by the considered algorithm is argument-bounded. The cases of the difference algorithm are generated in parallel to the proof steps of the meta-induction. We illustrate the instance test and the synthesis of the difference algorithm with the algorithm function \( \text{remove}(n : \text{number} x : \text{list}) : \text{list} \leftarrow \cdots \) from Section 1.

For performing the instance test, i.e. to verify whether \( \text{remove} \) computes a 2-bounded function, the representative \([r]\) of the result term of each case "if \( \varphi \) then \( r \)" is compared with the representative \([x]\) of the second formal parameter,
i.e. $\vdash_{L} \langle [r] \leq_{\#} [x], \Delta \rangle$ is computed. We find $\vdash_{L} \langle \text{empty} \leq_{\#} \text{empty}, \text{false} \rangle$ for the non-recursive case "if $x = \text{empty}$ then $x$" which means that $\text{remove}$ is 2-bounded at least for this case. Since the difference equivalent is computed as $\text{false}$, the result in the non-recursive case never is $\leq_{\#}$-smaller than the input argument $x$. Therefore the case "if $x = \text{empty}$ then false" is inserted into the definition of the 2-difference algorithm function $\Delta^{2}\text{remove}(n : \text{number} x : \text{list}) : \text{bool} \leq \cdots$ under construction. This completes the base case of the meta-induction.

In each recursive case we assume

$$\text{true} \langle \text{remove}(n \text{ tail}(x)) \leq_{\#} \text{tail}(x), \Delta^{2}\text{remove}(n \text{ tail}(x)) \rangle$$

as the induction hypothesis. Now if $\vdash_{L} \langle \text{tail}(x) \leq_{\#} t, \Delta \rangle$ for a term $t$, then $\text{true} \langle \text{tail}(x) \leq_{\#} t, \Delta \rangle$ by the soundness of the E-calculus. Hence $\text{true} \langle \text{remove}(n \text{ tail}(x)) \leq_{\#} t, \Delta^{2}\text{remove}(n \text{ tail}(x)) \lor \Delta \rangle$ by the induction hypothesis. Therefore the E-calculus remains sound if we add an estimation rule

$$\frac{\langle \text{tail}(x) \leq_{\#} t, \Delta \rangle}{\langle \text{remove}(n \text{ tail}(x)) \leq_{\#} t, \Delta^{2}\text{remove}(n \text{ tail}(x)) \lor \Delta \rangle} \quad \text{for all } t, \Delta \quad (6.1)$$

to the rules of the E-calculus. The E-calculus also remains decidable by the same argumentation as used in Section 4.6. The additional rule (6.1) is an instance of the argument estimation rule for $\text{remove}$, viz.

$$\frac{\langle t_{2} \leq_{\#} t_{1}, \Delta \rangle}{\langle \text{remove}(t_{1} t_{2}) \leq_{\#} t_{1}, \Delta^{2}\text{remove}(t_{1} t_{2}) \lor \Delta \rangle} \quad \text{for all } t_{1}, t_{2}, t, \Delta \quad (6.2)$$

because (6.1) can be obtained from (6.2) by substituting $n$ for $t_{1}$ and $\text{tail}(x)$ for $t_{2}$. Since we are in a proof for verification that $\text{remove}$ is 2-bounded, i.e. that (6.2) is a sound estimation rule, (6.2) cannot be used in the proof. However the additional rule (6.1) can be used by an inductive argument.

We let $\vdash_{L+}$ denote a derivation in the E-calculus which is extended by (6.1), and then $\vdash_{L+} \langle \text{remove}(n \text{ tail}(x)) \leq_{\#} \text{add(head(x)} \text{ tail}(x)), \text{true} \rangle$ is computed for the first recursive case "if $x \not= \text{empty} \land \text{head}(x) = n \text{ then } \text{remove}(n \text{ tail}(x))". This means that $\text{remove}$ is 2-bounded also for this case. Since the difference equivalent is computed as $\text{true}$, the result in the first recursive case always is $\leq_{\#}$-smaller than the input argument $x$. Therefore the case "if $x \not= \text{empty} \land \text{head}(x) = n \text{ then true}" is inserted into the definition of the 2-difference algorithm function $\Delta^{2}\text{remove}$ under construction. This completes the first step case of the meta-induction.

\footnote{For sake of readability we do not present each algorithm in psnf (cf. Section 5.3). Instead we "treat" all algorithms as if they were given in psnf. For instance, $x \not= \text{empty}$ has to be read as $x = \text{add(head(x) tail(x))}$ and then a nontrivial representative for $x$ can be computed.}
Finally we compute \( \sim r \div (\text{add}(\text{head}(x) \text{ remove}(n \text{ tail}(x)))) \) for the second recursive case "if \( x \neq \text{empty} \land \text{head}(x) \neq n \) then \( \text{add}(\text{head}(x) \text{ remove}(n \text{ tail}(x))) \)”, where also here (6.1) is used as an additional estimation rule which soundness is guaranteed by the induction hypothesis. Hence \( \text{remove} \) is 2-bounded also for this case. Since the difference equivalent is computed as \( \Delta^2\text{remove}(n \text{ tail}(x)) \), the result in the second recursive case is \(<_\varphi\)-smaller than the input argument \( x \) iff \( \Delta^2\text{remove}(n \text{ tail}(x)) \) is satisfied. Therefore the case “if \( x \neq \text{empty} \land \text{head}(x) \neq n \) then \( \Delta^2\text{remove}(n \text{ tail}(x)) \)” is inserted into the definition of the 2-difference algorithm \( \Delta^2\text{remove} \) under construction. This completes the second step case of the meta-induction and the whole proof.

Summing up, \( \text{remove} \) is recognized as a 2-bounded function with 2-difference function \( \Delta^2\text{remove} \) (usually known as \textit{member}), where this difference function is computed by a 2-difference algorithm synthesized as

\[
\text{function } \Delta^2\text{remove}(n : \text{number} \ x : \text{list}) : \text{bool} \leftarrow \\
\quad \text{if } x = \text{empty} \text{ then false} \\
\quad \text{if } x \neq \text{empty} \land \text{head}(x) = n \text{ then true} \\
\quad \text{if } x \neq \text{empty} \land \text{head}(x) \neq n \text{ then } \Delta^2\text{remove}(n \text{ tail}(x)).
\]

6.2. Recognizing argument-bounded algorithms

Analyzing the above deduction for establishing the 2-boundedness of \( \text{remove} \) we see that all we needed was induction (based on the recursions in the algorithm) and the estimation relation (and recursion and the difference equivalent for the synthesis of the difference algorithm). Therefore we can formulate decidable syntactical requirements for an algorithm which imply that the computed function is argument-bounded.

This approach generally works because algorithms which compute \( p \)-bounded functions often have a \textit{similar shape and structure} as the algorithm \( \text{remove} \) and the subsequently presented algorithms for \textit{minus2}, \textit{half2} and \textit{log2} have. This is an empirical observation which was obtained by the analysis of several algorithms (cf. [23]). The similarity of these algorithms can be formally expressed by an algorithm schema for \( p \)-bounded algorithms such that each instance of this schema computes a \( p \)-bounded function:

**Definition 6.1.** An algorithm function \( g(x_1 : s_1 \ldots x_n : s_n) : s \leftarrow \cdots \) in an admissible specification \( S \) is \( p \)-bounded for some \( p \in \{1, \ldots, n\} \) iff for each case \"if \( \varphi \) then \( r \)\" of the algorithm

\[
\vdash_{\Gamma(S)} \langle [r]_{\varphi} \leq_{\#} [x_p]_{\varphi}, \gamma \rangle \quad \text{for some formula } \gamma,
\]

where \( \vdash_{\Gamma(S)} \) denotes derivability in the E-calculus which is extended by additional estimation rules.
for each recursive call \( g(t_1, \ldots, t_n) \) in \( r \). An algorithm is **argument-bounded** iff it is \( p \)-bounded for some argument position \( p \).

The \( p \)-difference algorithm function \( \Delta^p g(x_1 : s_1, \ldots, x_n : s_n) : \text{bool} \leftarrow \cdots \) for a \( p \)-bounded algorithm function \( g(x_1 : s_1, \ldots, x_n : s_n) : s \leftarrow \cdots \) contains exactly a pair of cases

\[
\begin{align*}
\text{if } \varphi \land \gamma & \text{ then true,} \\
\text{if } \varphi \land \neg \gamma & \text{ then false,}
\end{align*}
\]

for each case "if \( \varphi \) then \( r \)" in \( g \), where \( \gamma \) is given as in (6.3).

Derivability \( \vdash_{\Gamma(S)^+} \) in the extended E-calculus depends on the cases of the algorithm which is tested for \( p \)-boundedness (cf. (6.4)). \( \vdash_{\Gamma(S)^+} \) coincides with \( \vdash_{\Gamma(S)} \) for all non-recursive cases because no additional estimation rules exist, and \( \vdash_{\Gamma(S)^+} \) varies with the recursive calls in each recursive case.

The algorithm schema of Definition 6.1 is **decidable**, because (6.3) is the only requirement and \( \vdash_{\Gamma(S)^+} \) is decidable (cf. Section 4.6). Consequently the synthesis of a difference algorithm is **recursive**. Each \( p \)-difference algorithm is deterministic and case-complete "by construction", because each \( p \)-bounded algorithm is. Since only a terminating algorithm is tested for \( p \)-boundedness, its associated \( p \)-difference algorithm also terminates "by construction" because it uses a recursion \( \Delta^p g(t_1, \ldots, t_n) \) as a subformula of \( \gamma \) iff the \( p \)-bounded algorithm uses a recursion \( g(t_1, \ldots, t_n) \) in the corresponding case. Consequently an admissible specification remains admissible if it is extended by a \( p \)-difference algorithm.

The syntactical requirement demanded for an argument-bounded algorithm function \( g \) and the cases stipulated for its difference algorithm function \( \Delta^p g \) allow to prove that \( g \) is a \( p \)-bounded function and \( \Delta^p g \) is a \( p \)-difference function for \( g \) as expressed by the following theorem (see Appendix B for the proof):
if $x \neq \text{empty} \land \text{head}(x) = n \land \neg \text{TRUE}$ then false
if $x \neq \text{empty} \land \text{head}(x) \neq n \land \Delta^2\text{remove}(\text{tail}(x))$ then true
if $x \neq \text{empty} \land \text{head}(x) \neq n \land \neg \Delta^2\text{remove}(\text{tail}(x))$ then false.

This algorithm differs from the 2-difference algorithm presented in Section 6.1, but it should be obvious how trivial transformations and simplifications yield the version of function $\Delta^2\text{remove}$ as presented before. To ease readability we subsequently present examples of difference algorithms only in their simplified form (see Section 6.3 for some simplification techniques).

We give some examples for $p$-bounded algorithms. The following algorithm

```plaintext
function minus2(n, m : number) : number <=
  if m = 0 then n
  if m \neq 0 then minus2(pred(n) pred(m))
```

is 1-bounded because

\[
\begin{align*}
\vdash & \langle n \leq_* n, \text{FALSE} \rangle, \\
\vdash & \langle \text{minus2}(\text{pred}(n) \text{pred}(m)) \leq_* n \rangle, \\
& \Delta^1\text{minus2}(\text{pred}(n) \text{pred}(m)) \lor n \neq 0 \rangle,
\end{align*}
\]

where $n \neq 0$ abbreviates $\Delta^1\text{pred}(n)$ and

\[
\frac{\langle \text{pred}(n) \leq_* t, \Delta \rangle}{\langle \text{minus2}(\text{pred}(n) \text{pred}(m)) \leq_* t, \Delta^1\text{minus2}(\text{pred}(n) \text{pred}(m)) \lor \Delta \rangle}
\]

for all $t, \Delta$

is used as an additional estimation rule. Therefore the (simplified) 1-difference algorithm for $\text{minus2}$ is synthesized as

```plaintext
function \Delta^1\text{minus2}(n, m : number) : bool <=
  if m = 0 then false
  if m \neq 0 \land n \neq 0 then true
  if m \neq 0 \land n = 0 then \Delta^1\text{minus2}(\text{pred}(n) \text{pred}(m)).
```

The following algorithm

```plaintext
function half2(n : number) : number <=
  if n = 0 then 0
  if n \neq 0 \land \text{pred}(n) = 0 then 0
  if n \neq 0 \land \text{pred}(n) \neq 0 then \text{succ}(\text{half2}(\text{pred}(\text{pred}(n))))
```

is 1-bounded because
\[ \vdash_r \langle 0 \equiv_0 0, \text{FALSE} \rangle , \]
\[ \vdash_r \langle 0 \equiv_0 \text{succ}(0), \text{TRUE} \rangle , \]
\[ \vdash_{r^+} \langle \text{succ}(\text{half2}(\text{pred}(\text{pred}(n)))) \]
\[ \equiv_0 \text{succ}(\text{succ}(\text{pred}(\text{pred}(n)))) , \text{TRUE} \rangle , \]

where

\[ \langle \text{pred}(\text{pred}(n)) \equiv_0 t, \Delta \rangle \]
\[ \vdash \langle \text{half2}(\text{pred}(\text{pred}(n))) \equiv_0 t, \Delta^1 \text{half2}(\text{pred}(\text{pred}(n))) \lor \Delta \rangle \]

for all \( t, \Delta \)

is used as an additional estimation rule. Therefore the (simplified) 1-difference algorithm for \text{half2} is synthesized as

\begin{verbatim}
function \Delta^1 \text{half2}(n : number) : bool <=
   if n = 0 then false
   if n \neq 0 \land \text{pred}(n) = 0 then true
   if n \neq 0 \land \text{pred}(n) \neq 0 then true .
\end{verbatim}

The following algorithm

\begin{verbatim}
function \text{log2}(n : number) : number <=
   if n = 0 then 0
   if n \neq 0 \land \text{pred}(n) = 0 then 0
   if n \neq 0 \land \text{pred}(n) \neq 0 then
      \text{succ(}\text{log2(}\text{succ(}\text{half2(}\text{pred(}\text{pred(}n)\text{)})\text{)})\text{)})
\end{verbatim}

is 1-bounded because

\[ \vdash_r \langle 0 \equiv_0 0, \text{FALSE} \rangle , \]
\[ \vdash_r \langle 0 \equiv_0 \text{succ}(0), \text{TRUE} \rangle , \]
\[ \vdash_{r^+} \langle \text{succ}(\text{log2}(\text{succ}(\text{half2}(\text{pred}(\text{pred}(n))))))) \equiv_0 \text{succ}(\text{pred}(\text{pred}(n)))) , \]
\[ \Delta^1 \text{log2}(\text{succ}(\text{half2}(\text{pred}(\text{pred}(n)))))) \lor \text{pred}(\text{pred}(n)) \neq 0 \rangle , \]

where the 1-boundedness of \text{half2} is required to verify the 1-boundedness of \text{log2}, i.e. we assume \text{half2} \in \Gamma_1, \text{pred}(\text{pred}(n)) \neq 0 abbreviates \Delta^1 \text{half2}(\text{pred}(\text{pred}(n))) and

\[ \langle \text{succ}(\text{half2}(\text{pred}(\text{pred}(n)))) \equiv_0 t, \Delta \rangle \]
\[ \vdash \langle \text{log2}(\text{succ}(\text{half2}(\text{pred}(\text{pred}(n))))))) \equiv_0 t , \Delta^1 \text{log2}(\text{succ}(\text{half2}(\text{pred}(\text{pred}(n)))))) \lor \Delta \rangle \]

for all \( t, \Delta \)
is used as an additional estimation rule. Therefore the (simplified) 1-difference algorithm for \( \log_2 \) is synthesized as

\[
\begin{align*}
\text{function } & \Delta^1 \log_2(n : \text{number}) : \text{bool} \leftarrow \\
& \text{if } n = 0 \text{ then false} \\
& \text{if } n \neq 0 \land \text{pred}(n) = 0 \text{ then true} \\
& \text{if } n \neq 0 \land \text{pred}(n) \neq 0 \land \text{pred}(\text{pred}(n)) \neq 0 \text{ then true} \\
& \text{if } n \neq 0 \land \text{pred}(n) \neq 0 \land \text{pred}(\text{pred}(n)) = 0 \text{ then} \\
& \Delta^1 \log_2(\text{succ}(\text{half}_2(\text{pred}(\text{pred}(n))))).
\end{align*}
\]

6.3. Optimization of difference algorithms

Since the definition formulas of a difference algorithm are used in termination proofs (cf. Section 2.3), we are interested in obtaining a difference algorithm as simple as possible. This is of proof-technical relevance because the simpler the difference algorithms are, the easier the termination hypotheses are to prove.

The synthesized difference algorithms often contain redundant case-conditions, superfluous cases and unnecessary recursions. For instance, the 2-difference algorithm \( \Delta^2 \text{remove} \) given in Section 6.2 has redundant case-conditions of the form \( \varphi \land \neg \text{FALSE} \) and \( \varphi \land \text{TRUE} \) which can be simplified. It also contains superfluous cases of the form "if... \land \text{FALSE} then..." and "if... \land \neg \text{TRUE} then..." which can be eliminated. Such modifications of a difference algorithm are performed by condition subsumption which uses an induction theorem prover to verify \([\forall \cdots \varphi \rightarrow \gamma] \in Th(S)\) for each pair of cases

\[
\begin{align*}
& \text{if } \varphi \land \gamma \text{ then true} , \\
& \text{if } \varphi \land \neg \gamma \text{ then false} ,
\end{align*}
\]

in a difference algorithm. If successful both cases are replaced by "if \( \varphi \) then true". Otherwise it is tested for \([\forall \cdots \varphi \rightarrow \neg \gamma] \in Th(S)\), and "if \( \varphi \) then false" replaces both cases if this test succeeds.

Difference algorithms are also modified by case merging, which means that each pair of cases

\[
\begin{align*}
& \text{if } \varphi \land L \text{ then } r , \\
& \text{if } \varphi \land \neg L \text{ then } r ,
\end{align*}
\]

in a difference algorithm is replaced by "if \( \varphi \) then \( r \)". For instance, case merging simplifies the 1-difference algorithm for \( \text{half}_2 \) from Section 6.2 to

\[
\begin{align*}
\text{function } & \Delta^1 \text{half}_2(n : \text{number}) : \text{bool} \leftarrow \\
& \text{if } n = 0 \text{ then false} \\
& \text{if } n \neq 0 \text{ then true} .
\end{align*}
\]

The costs of a termination proof (measured in the number of inductive
argumentations) directly depends on the number of recursive calls in a difference algorithm and therefore it is useful to eliminate recursions in a difference algorithm if possible.\footnote{function $\Delta^r$remove is an example of a difference algorithm which cannot be defined without recursion.} For instance, the difference algorithm $\Delta^1\log 2$ from Section 6.2 contains a recursive case

\[
\text{if } n \neq 0 \land \text{pred}(n) \neq 0 \land \text{pred}(\text{pred}(n)) = 0 \text{ then }
\]

\[
\Delta^1\log 2(\text{succ}(\text{half}^2(\text{pred}(\text{pred}(n)))))) \tag{6.5}
\]

and consequently the proof of a termination hypothesis as $[\forall n: \text{number } n \neq 0 \rightarrow \Delta^1\log 2(n)]$ requires induction. After replacement of the recursive call by true, case merging can be applied which yields the simplified difference algorithm

\[
\text{function } \Delta^1\log 2(n: \text{number}): \text{bool} \Leftarrow \\
\text{if } n = 0 \text{ then false} \\
\text{if } n \neq 0 \text{ then true.}
\]

Now $[\forall n: \text{number } n \neq 0 \rightarrow \Delta^1\log 2(n)]$ can be proved by propositional reasoning only. We therefore always attempt to eliminate recursions in a difference algorithm function $\Delta^p g(x_1 : S_1 \ldots x_n : S_n): \text{bool} \Leftarrow \cdots$ by replacing some recursive cases with non-recursive cases.

Let $\Psi$ be some non-empty set of formulas such that “if $\psi_i$ then $\Delta^p g(t_{i,1} \ldots t_{i,n})$” is a case in function $\Delta^p g$ for each $\psi_i \in \Psi$ and let $\Phi_b$ be the set of all formulas $\varphi_i$ such that “if $\varphi_i$ then $b$” is a case in function $\Delta^p g$ where $b \in \{\text{true, false}\}$. Now if

\[
[\forall x^*: w \psi_i \rightarrow \bigvee_{\varphi \in \Psi} \delta_i(\psi) \lor \bigvee_{\varphi \in \Phi_b} \delta_i(\varphi)] \in T h(S) \tag{6.6}
\]

can be verified for each $\psi_i \in \Psi$, where $\delta_i = \{x_1/t_i,1, \ldots, x_i/t_i,n\}$, then each recursive case “if $\psi_i$ then $\Delta^p g(t_{i,1} \ldots t_{i,n})$” can be replaced by “if $\psi_i$ then $b$” without altering the operation computed by function $\Delta^p g$. The formula in (6.6) is a so-called recursion elimination formula and the truth of all these formulas guarantees that an equivalent difference algorithm is obtained if recursive calls are eliminated.

Assume that some $\psi_h \in \Psi$ is satisfied for an input $q_0^*$ of function $\Delta^p g$. Since function $\Delta^p g$ terminates, $q_0^*$ necessitates finitely many, say $k$, successive recursive calls with the cases given by $\Psi$. Hence some $\psi_i$ holds on the $k$th call for the actual input $q_k^*$ but no $\delta_i(\psi)$ holds for $q_k^*$, because otherwise we would have $k + 1$ successive recursive calls with the cases given by $\Psi$. Hence with (6.6) some $\delta_i(\psi)$ is satisfied for $q_k^*$ and function $\Delta^p g$ returns $\delta_i(b) = b$, i.e. the same result which is computed by the modified difference algorithm. If no $\psi_h \in \Psi$ is satisfied for $q_0^*$, then some other condition $\gamma$ of a case “if $\gamma$ then $r$” is satisfied for $q_0^*$. Since $\gamma \not\in \Psi$ this case is also contained in the modified algorithm, and therefore both algorithms return the same result.

For the difference algorithm $\Delta^1\log 2$ from Section 6.2 we find $\Psi = \{n \neq 0 \land \text{pred}(n) \neq 0 \land \text{pred}(\text{pred}(n)) = 0\}$ and with $\Phi_{\text{true}} = \{n \neq 0 \land \text{pred}(n) = 0, n \neq 0 \land \text{pred}(n) \neq 0 \land \text{pred}(\text{pred}(n)) \neq 0\}$ the recursion elimination formula
\( \forall n : \text{number} \)
\[
\left( n \neq 0 \land \text{pred}(n) \neq 0 \land \text{pred}(\text{pred}(n)) = 0 \rightarrow \ight.
\]
\[
\text{half2}(\text{pred}(\text{pred}(n))) \neq 0 \land \\
\text{pred}(\text{half2}(\text{pred}(\text{pred}(n)))) = 0 \lor \\
\text{half2}(\text{pred}(\text{pred}(n))) = 0 \lor \\
\text{half2}(\text{pred}(\text{pred}(n))) \neq 0 \land \\
\text{pred}(\text{half2}(\text{pred}(\text{pred}(n)))) \neq 0 \\
\]
(6.7)

is obtained, where we have already performed trivial simplifications like replacement of \( \text{succ}(\ldots) \neq 0 \) by \( \text{TRUE} \) and \( \text{pred}(\text{succ}(t)) \) by \( t \) to ease readability. This recursion elimination formula can be easily proved by propositional reasoning and this justifies the soundness of the recursion elimination in function \( \Delta^1 \text{log2} \).

Also the recursion in the difference algorithm \( \Delta^1 \text{minus2} \) of Section 6.2 can be replaced by \( \text{false} \) after verification of the recursion elimination formula

\( \forall n, m : \text{number} \)
\[
\left( m \neq 0 \land n = 0 \rightarrow \text{pred}(m) \neq 0 \land \text{pred}(n) = 0 \lor \text{pred}(m) = 0 \right) . \\
(6.8)

7. When termination cannot be proved

Of course there are cases for which our method fails: Termination cannot be verified for non-strongly terminating algorithms as McCarthy’s 91-function (cf. Section 3) and we did not consider mutual recursion. Also termination cannot be verified if a well-founded relation which is not based on the \( s \)-size measure (cf. Section 4.1) is required for the termination proof. For instance, our method is unable to prove termination of the algorithm

\textbf{function} \( \text{subtract1}(n, m : \text{number}) : \text{number} \leftarrow \)
\[
\text{if } n \leq m \text{ then } 0 \\
\text{if } n > m \text{ then } \text{succ}(\text{subtract1}(n \text{ succ}(m))) .
\]

Using \( \tau(n, m) := n - m \) as a termination function, termination of \( \text{subtract1} \) is easily verified but this termination argument cannot be expressed by estimation formulas.

Our method also fails if an argument-bounded function \( g \) is defined by an algorithm which is not argument-bounded because then a termination hypothesis cannot be generated for an algorithm which “uses” \( g \) in a recursive call. For instance,

\[\vdots\]
\[\begin{array}{c}
16\quad \text{Sometimes, as for } \Delta^1 \text{log2, recursions can also be eliminated by symbolic evaluation. But this technique is too weak in general, because if fails e.g. for the recursion elimination in } \Delta^1 \text{minus2.}
\end{array}\]
function reverse(x : list) : list ≡
    if x = empty then x
    if x ≠ empty then append(reverse(tail(x)) add(head(x) empty))

defines a 1-bounded function because TRUE(reverse(x) ∼ₘ x, FALSE), but the
algorithm is not 1-bounded because append(reverse(tail(x)) add(head(x)
empty)) ∉ₘ add(head(x) tail(x)). Therefore reverse ∉ Γ₁ and the termination
hypothesis [∀x : list x ≠ empty → TRUE] cannot be generated for

function shuffle(x : list) : list ≡
    if x = empty then x
    if x ≠ empty then add(head(x) shuffle(reverse(tail(x)))) .

This problem is caused by the incompleteness of the E-calculus but (by our
experience) such failures rarely occur in practice because algorithms like reverse
are rarely used in the recursive calls of other algorithms.

Another example is the algorithm

function mod(n, m : number) : number ≡
    if m = 0 then m
    if n < m then n
    if n ≥ m ∧ m ≠ 0 then mod(minus(n m) m) ,

which defines a 2-bounded function because TRUE(mod(n m) ∼ₘ m, m ≠ 0). This
algorithm is 1-bounded but it is not 2-bounded because n ∉ₘ m. Here a remedy is
to replace the result n in the second case of mod by min(n m), where min
(computing the minimum of its inputs) is defined by a 1- and 2-bounded
algorithm. Since min(n m) ∼ₘ n the modified algorithm still is 1-bounded and with
min(n m) ∼ₘ m the algorithm now is also 2-bounded. This solution seems quite ad
hoc but (besides mod) we only know of one (non-artificial) algorithm, viz.
greatest.factor in [2], which requires such a modification.

Sometimes argument-boundedness cannot be verified because a result term in
an algorithm is not “maximally evaluated”. For instance, the 1-boundedness of
log (cf. Section 5.2) cannot be verified because half(succ(succ(pred(pred
(n)))))) ∉ₘ succ(pred(pred(n))). Such problems are avoided if terms are evaluated
before they are compared by the estimation relation. Here half(succ(pred-
(pred(n)))) must be replaced by succ(half(pred(pred(n)))) and then 1-bounded-
ness can be verified, cf. log2 in Section 6.2.

The test for argument-boundedness also fails if a required term representative
cannot be computed because the necessary structure equation is only implicitly
given. We have already demanded that algorithms are in positive structure normal
form for avoiding these problems (cf. Section 5.3), but this is not a general
remedy. For instance,
function \textit{subtract2}(n, m : number) : number <=

\textbf{if} n \leq m \textbf{then} 0

\textbf{if} n > m \textbf{then} \textit{succ}(\textit{subtract2}(\textit{pred}(n), m))

defines a 1-bounded function because \textit{TRUE} \left( \textit{subtract2}(n, m) \leq n, n \neq 0 \land m \neq 0 \right), but the algorithm is not 1-bounded since \textit{succ}(\textit{subtract2}(\textit{pred}(n), m)) \not\leq n \land n = [n].

Here the condition in the recursive case must be replaced by \textit{n > m} \land n = \textit{succ}(\textit{pred}(n)) and then 1-boundedness can be verified since now $[n] = \textit{succ}(\textit{pred}(n))$.

Finally, the method fails if results are stipulated for “don’t care” inputs such that the function defined by an algorithm is not argument-bounded only for this reason. For instance, the algorithm \textit{gcd} from Section 4.1 still terminates if the first case in \textit{mod} is given as “\textbf{if} m = 0 \textbf{then} \textit{succ}(0)”.

But then \textit{mod} is not 1-bounded because \textit{succ}(0) \not\leq n and consequently no termination hypothesis can be generated for \textit{gcd}. Therefore “helpful” results, as 0, should be stipulated for irrelevant function applications, as \textit{mod}(n, 0), such that argument-boundedness is supported (cf. Section 4.2).

8. Related work

Research on proving termination is closely related to research on program verification in general. The idea of using termination functions and properties of well-founded sets for proving termination of flow chart programs is suggested by Floyd [7].

Manna [10] presents a procedure which synthesizes a first-order formula $\varphi_P$ from a flow chart program $P$ such that $P$ terminates under an interpretation $I$ (which assigns operations over particular domains to the operation symbols in $P$) if and only if $\varphi_P$ is valid under the same interpretation. For proving the formula $\varphi_P$ axioms $\Phi_P$ which are satisfied by $I$ and also axioms modeling the involved operations are needed.\footnote{Note that $\varphi_P$ as well as $\Phi_P$ contain “new” relation symbols not in $P$.} The axioms $\Phi_P$ define a well-founded relation by which $P$ terminates or else contradict $\varphi_P$. Consequently the problem of finding a termination argument for $P$ is mapped into the problem of finding the adequate axiom set $\Phi_P$. Viewed in this context, our method computes an axiom set $\Phi_P$ such that $\Phi_P$ is satisfied by $I$ and a proof of $\varphi_P$ from $\Phi_P$ can be uniformly obtained. For the introductory example from Section 1, for instance, $\Phi_P$ contains only formula (1.5). Our procedure fails in the computation of $\Phi_P$ if $P$ does not terminate, but it may fail also for programs which do terminate.

A system for automated program verification is presented in [5], where proving termination is essentially based on Floyd’s ideas. The paper describes a semi-automatic facility for the generation and verification of convergence conditions which are sufficient for the termination of the loops in a program.
Based on Hoare's method, Manna and Pnueli [13] present an axiomatic approach which allows to prove both correctness and termination of while-programs by one unified formalism. Here termination of a while-statement is proved by defining a termination function (called convergence function in the paper) and then to infer with the proposed inference rules that the value of the termination function decreases by the execution of the loop. Examples are given, including an algorithm for the greatest common divisor and the partition problem from Hoare's FIND algorithm [8]. The example proofs in the paper give an impression of the complexity and difficulty of verifying termination with a rigorous formal statement.  

Katz and Manna [9] compare four methods for proving termination. Floyd's technique [7] and the loop approach use termination functions, where in the latter approach an upper bound for each increasing counter of a loop has to be established. In the exit approach termination is proved by showing that some so-called exit conditions, which have to be generated for a loop, are satisfied at some stage of the computation. Unlike the former methods (and our proposal), also non-termination can be handled with this method. Finally Burstall's method [4] which proves termination and correctness simultaneously by structural induction is discussed. The success of this method depends on the invention of the "right" well-founded order for the given verification problem. Burstall's method seems especially suited to show the termination of algorithms using nested recursions and therefore being difficult to understand and to verify (cf. Section 3). Because one needs to know what these algorithms do when proving termination, correctness and termination have to be shown simultaneously. This observation is also well recognized in the papers of Boyer and Moore [3] and of Paulson [17] who compares three methods for proving termination, viz. using measurement functions, in domain theory using LCF and showing that the relation defined by the recursive calls is well-founded. Burstall's approach is further developed with the intermittent-assertion method [14].

A system for proving termination of PROLOG programs is presented in [1] where termination properties are proved by reasoning with specific equations which have to be associated with a program. Proving termination of PROLOG programs by inferring certain inequations is considered in [21] and in [18].

Our proposal for proving termination has similarities with the method implemented in the NQTHM system of Boyer and Moore [2]. There, an initial specification $S_0$ is assumed which contains the data structures $bool$ and $number$ and the algorithms $plus$ and $lt$ (cf. Section 2). Each data structure $s$ implicitly

---

*The recursive versions of the algorithms for the greatest common divisor and the partition problem can be proved to be terminating with our method, cf. [23]. It must be admitted, however, that proving termination is much more complicated for algorithms defined in an imperative language like ALGOL 60 with assignments, goto- and while-statements than for algorithms given in a pure functional notation.*
defines an algorithm function \( \text{count}_s(x:s) : \text{number} \leq \cdots \) which provides an algorithmic definition of the \( s \)-size measure \( #_s \) as defined in Section 4.1. Therefore each estimation formula \( \langle t \preceq r, \Delta \rangle \) could be represented in the NQTHM system by the first-order formulas \( \forall \cdot \cdot \cdot \Delta \leftrightarrow \text{lt}(\text{count}_s(t) \text{count}_s(r)) \) and \( \forall \cdot \cdot \cdot \Delta \leftrightarrow \text{lt}(\text{count}_s(t) \text{count}_s(r)) \). For proving termination so-called induction lemmata are formulated by the user of the system. These formulas must be proved by the system and if successful they are associated with a certain label such that the system remembers them when proving termination. The formula

\[
\forall n : \text{number} \forall x : \text{list} \\
\quad \text{member}(n \ x) \rightarrow \text{count_list}(\text{remove}(n \ x)) < \text{count_list}(x)
\]  

is an example of an induction lemma for \( \text{remove} \), where \( t < r \) abbreviates \( \text{lt}(t \ r) \), and an induction lemma

\[
\forall x_1 : s_1 \ldots x_n : s_n \ \Delta^p g(x_1 \ldots x_n) \rightarrow \text{count}_s(g(x_1 \ldots x_n)) < \text{count}_s(x_p)
\]  

(8.2)

can be formulated for each \( p \)-bounded algorithm \( g \) with difference algorithm \( \Delta^p g \) by the user of the NQTHM system. Now if termination of an algorithm like \( \text{sort} \) from Section 1 has to be proven, the system recognizes the call of \( \text{remove} \) in the recursion of \( \text{sort} \), remembers the induction lemma for \( \text{remove} \) and generates the same termination hypothesis for \( \text{sort} \), viz. (1.4), as our method does.

The induction lemma technique is more powerful than our method because measurement algorithms other than \( \text{count}_s \) may be used in the formulation of an induction lemma. For instance, a termination hypothesis for the algorithm \( \text{subtract1} \) from Section 7 can be computed with the induction lemma

\[
\forall x, y : \text{number} \quad x < y \rightarrow \text{minus}(y \ \text{succ}(x)) < \text{minus}(y \ x).
\]  

(8.3)

This means that strong termination can be proved also for algorithms for which our method fails.

The induction lemma technique provides less automatization than our proposal because the measurement algorithms and the induction lemmata have to be formulated by the user of the system. This means that the idea why an algorithm terminates has to be found by a human. Our method also requires less proof obligations. For instance, for proving the termination of \( \text{sort} \) in the NQTHM system the user has to define the algorithm \( \text{member} \) and to formulate the induction lemma (8.1), which is performed automatically by the synthesis of a difference algorithm in our method. Then the system has to prove the termination of \( \text{member} \) and the truth of the induction lemma, which is guaranteed "by construction" in our method. The remaining steps, viz. the generation and the proof of the termination hypothesis for \( \text{sort} \), are similar in both methods.
9. Conclusion

Our method has been implemented in the induction theorem prover INKA.\textsuperscript{19} After a new data structure is defined, the system labels each reflexive selector as 1-bounded and computes the corresponding 1-difference algorithm as described in Section 6.1. If a new algorithm is defined, termination hypotheses are generated for the algorithm as discussed in Section 5, and the system tries to prove these conjectures. If successful, the algorithm is inserted into the data base. Then the system decides for each argument position $p$ whether the new algorithm is $p$-bounded and a $p$-difference algorithm is synthesized for each such argument position (cf. Section 6.2). After each $p$-difference algorithm has been optimized by condition subsumption, case merging and recursion elimination (cf. Section 6.3), it is also inserted into the data base and the system is ready for the next input.

Since this approach may also produce false conjectures some precautions must be taken to avoid wasting theorem proving resources. If an algorithm does not terminate, our method computes a false termination hypothesis (if it computes a termination hypothesis at all). Also an optimization may not be applicable and then false formulas are generated by condition subsumption or by recursion elimination. Therefore the system tries to falsify each system-computed conjecture before a proof attempt is made (cf. [19]).

The INKA system performs successfully on many algorithms. A collection of 50 algorithms is presented in [23], including classical sorting algorithms and algorithms for standard arithmetical operations, and it is shown which termination hypotheses are generated for the algorithms in the collection. Generally termination proofs require induction, but—as it can be seen from the examples—the generated termination hypotheses often can be proved by cases and propositional reasoning only (cf. Appendix A).

The data base from [2] was also used as a benchmark. In this data base 73 algorithms terminate with selector functions and 6 terminate with other argument-bounded algorithms. Our system proves the termination of all these algorithms automatically (where the algorithm for $\text{greatest.factor}$ must be modified as described in Section 8). This means that our system implicitly synthesizes all the induction lemmata which have to be submitted to the NQTHM system by the system user. But three algorithms in this data base terminate with a well-founded relation which is not based on the $s$-size measure, viz. $\text{normalize}$, $\text{gopher}$ and $\text{samefringe}$, and consequently our method fails for these algorithms.

\textsuperscript{19} Induction theorem proving systems have to solve their own termination problems (and the work presented here was inspired by the development of such a system). Verification of termination is a central problem in automated induction because knowing the reason(s) why an algorithm terminates is the most important key to find a useful induction axiom for a formula (cf. [2, 24]).
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Appendix A

We illustrate our method by six sorting algorithms which are standard material in computer science ground courses. For each algorithm the termination hypotheses which are computed according to the termination criterion of Corollary 5.7 are presented. As may be observed from the presentation, all termination hypotheses can be proved by propositional reasoning and cases only.

We also present the difference algorithms (synthesized according to Definition 6.1) for the argument-bounded algorithms (recognized according to Definition 6.1) which are used in the sorting algorithms. All difference algorithms are given in their optimized form (cf. Section 6.3), and it is easily seen that all formulas justifying the soundness of the optimizations can be proved by propositional reasoning and cases only.

For saving space, auxiliary algorithms as function last(x : list) : number <= ... (for the computation of the rightmost element of a list) etc. are omitted if they do not illustrate our method.

A.1. Bubblesort

function bubble(x : list) : list <=
  if x = empty then x
  if x <> empty ^ tail(x) = empty then x
  if x <> empty ^ tail(x) <> empty ^ head(x) <= head(tail(x)) then
    add(head(tail(x)) bubble(add(head(x) tail(tail(x)))))
  if x <> empty ^ tail(x) <> empty ^ head(x) > head(tail(x)) then
    add(head(x) bubble(tail(x)))

computes a permutation of x such that the last list element is a minimal element of the list. bubble is a 1-bounded algorithm with 1-difference algorithm (obtained after recursion elimination)

function Δ1bubble(x : list) : bool <=
  if TRUE then false .
function \texttt{but.last}(x : list) : list \leftarrow \\
if \texttt{x} = \texttt{empty} then \texttt{empty} \\
if \texttt{x} \neq \texttt{empty} \land \texttt{tail(x)} = \texttt{empty} then \texttt{empty} \\
if \texttt{x} \neq \texttt{empty} \land \texttt{tail(x)} \neq \texttt{empty} then \texttt{add(head(x) but.last(tail(x)))}

computes a copy of \texttt{x} with the last list element removed. \texttt{but.last} is a 1-bounded algorithm with 1-difference algorithm (obtained after recursion elimination)

function \texttt{\Delta^1 but.last}(x : list) : bool \leftarrow \\
if \texttt{x} = \texttt{empty} then \texttt{false} \\
if \texttt{x} \neq \texttt{empty} then \texttt{true}

function \texttt{bubblesort}(x : list) : list \leftarrow \\
if \texttt{x} = \texttt{empty} then \texttt{empty} \\
if \texttt{x} \neq \texttt{empty} then \\
\texttt{add(last(bubble(x)) bubblesort(but.last(bubble(x))))}.

The termination hypothesis for \texttt{bubblesort} is computed as
\[ \forall x : \text{list} \quad x \neq \texttt{empty} \rightarrow \Delta^1 \texttt{but.last}(\texttt{bubble(x)}) \lor \Delta^1 \texttt{bubble(x)}. \]

A.2. Selection sort

function \texttt{replace}(n, m : number x : list) : list \leftarrow \\
if \texttt{x} = \texttt{empty} then \texttt{empty} \\
if \texttt{x} \neq \texttt{empty} \land \texttt{head(x)} = n then \texttt{add(m tail(x))} \\
if \texttt{x} \neq \texttt{empty} \land \texttt{head(x)} \neq n then \texttt{add(head(x) replace(n m tail(x)))}

replaces the leftmost occurrence of \texttt{n} in \texttt{x} with \texttt{m}. \texttt{replace} is a 3-bounded algorithm with 3-difference algorithm (obtained after recursion elimination)

function \texttt{\Delta^3 replace}(n, m : number x : list) : bool \leftarrow \texttt{if TRUE then false}

function \texttt{selectsort}(x : list) : list \leftarrow \\
if \texttt{x} = \texttt{empty} then \texttt{empty} \\
if \texttt{x} \neq \texttt{empty} \land \texttt{head(x)} = \texttt{minimum(x)} then \\
\texttt{add(head(x) selectsort(tail(x)))} \\
if \texttt{x} \neq \texttt{empty} \land \texttt{head(x)} \neq \texttt{minimum(x)} then \\
\texttt{add(minimum(x) selectsort(replace(minimum(x) head(x) tail(x))))}.

\[ \text{Bold italic symbols denote term representatives as here } x \text{ abbreviates } \text{add(head(x) tail(x))}. \]
The termination hypotheses for selectsort are computed as
\[ \forall x : \text{list} \quad x \neq \text{empty} \land \text{head}(x) = \text{minimum}(x) \rightarrow \text{TRUE} , \]
\[ \forall x : \text{list} \quad x \neq \text{empty} \land \text{head}(x) \neq \text{minimum}(x) \rightarrow \Delta^3 \text{replace}(\text{minimum}(x), \text{head}(x), \text{tail}(x)) \lor \text{TRUE} . \]

A.3. Minsort

function delete.minimum(x : list) : list \leftarrow
\begin{align*}
& \quad \text{if } x = \text{empty} \text{ then empty} \\
& \quad \text{if } x \neq \text{empty} \land \text{tail}(x) = \text{empty} \text{ then empty} \\
& \quad \text{if } x \neq \text{empty} \land \text{tail}(x) \neq \text{empty} \land \text{head}(x) \leq \text{head}(\text{tail}(x)) \text{ then} \\
& \qquad \quad \text{add}(\text{head}(\text{tail}(x)), \text{delete.minimum}(\text{add}(\text{head}(x) \text{ tail}(\text{tail}(x))))) \\
& \quad \text{if } x \neq \text{empty} \land \text{tail}(x) \neq \text{empty} \land \text{head}(x) > \text{head}(\text{tail}(x)) \text{ then} \\
& \qquad \quad \text{add}(\text{head}(x), \text{delete.minimum}(\text{tail}(x)))
\end{align*}
deletes a minimal element from \( x \) and returns a permutation of the remaining list. delete.minimum is a 1-bounded algorithm with 1-difference algorithm (obtained after recursion elimination)

function \( \Delta^1 \text{delete.minimum}(x : \text{list}) : \text{bool} \leftarrow \)
\begin{align*}
& \quad \text{if } x = \text{empty} \text{ then false} \\
& \quad \text{if } x \neq \text{empty} \text{ then true} .
\end{align*}

function minsort(x : list) : list \leftarrow
\begin{align*}
& \quad \text{if } x = \text{empty} \text{ then empty} \\
& \quad \text{if } x \neq \text{empty} \text{ then} \\
& \qquad \quad \text{add}(\text{minimum}(x), \text{minsort}(\text{delete.minimum}(x))) .
\end{align*}
The termination hypothesis for minsort is computed as
\[ \forall x : \text{list} \quad x \neq \text{empty} \rightarrow \Delta^1 \text{delete.minimum}(x) . \]

A.4. Quicksort

function smaller(n : number x : list) : list \leftarrow
\begin{align*}
& \quad \text{if } x = \text{empty} \text{ then empty} \\
& \quad \text{if } x \neq \text{empty} \land n < \text{head}(x) \text{ then } \text{smaller}(n, \text{tail}(x)) \\
& \quad \text{if } x \neq \text{empty} \land n \geq \text{head}(x) \text{ then } \text{add}(\text{head}(x), \text{smaller}(n, \text{tail}(x)))
\end{align*}
computes a copy of \( x \) with all list elements greater than \( n \) removed. smaller is a 2-bounded algorithm with 2-difference algorithm
function \( \Delta^2 \text{smaller}(n : \text{number} \ x : \text{list}) : \text{bool} \) 
\[
\begin{align*}
\text{if } x = \text{empty} & \text{ then false} \\
\text{if } x \neq \text{empty} \land n < \text{head}(x) & \text{ then true} \\
\text{if } x \neq \text{empty} \land n \geq \text{head}(x) & \text{ then } \Delta^2 \text{smaller}(n \ \text{tail}(x)),
\end{align*}
\]
i.e. \( \Delta^2 \text{smaller}(n \ x) \) returns \text{true} iff some element of \( x \) is greater than \( n \).

function \( \text{larger}(n : \text{number} \ x : \text{list}) : \text{list} \) 
\[
\begin{align*}
\text{if } x = \text{empty} & \text{ then } \text{empty} \\
\text{if } x \neq \text{empty} \land n < \text{head}(x) & \text{ then add(head(x) larger(n \ \text{tail}(x)))} \\
\text{if } x \neq \text{empty} \land n \geq \text{head}(x) & \text{ then larger(n \ \text{tail}(x))}
\end{align*}
\]
computes a copy of \( x \) with all list elements less than or equal \( n \) removed. \( \text{larger} \) is a \( \Delta^2 \)-bounded algorithm with \( \Delta^2 \)-difference algorithm.

function \( \Delta^2 \text{larger}(n : \text{number} \ x : \text{list}) : \text{bool} \) 
\[
\begin{align*}
\text{if } x = \text{empty} & \text{ then false} \\
\text{if } x \neq \text{empty} \land n < \text{head}(x) & \text{ then } \Delta^2 \text{larger}(n \ \text{tail}(x)) \\
\text{if } x \neq \text{empty} \land n \geq \text{head}(x) & \text{ then true},
\end{align*}
\]
i.e. \( \Delta^2 \text{larger}(n \ x) \) returns \text{true} iff some list element of \( x \) is less than or equal \( n \).

function \( \text{quicksort}(x : \text{list}) : \text{list} \) 
\[
\begin{align*}
\text{if } x = \text{empty} & \text{ then empty} \\
\text{if } x \neq \text{empty} & \text{ then append( quicksort(smaller(head(x) \ \text{tail}(x)) \\
& \quad \text{add(head(x) quicksort(larger(head(x) \ \text{tail}(x)))))}. 
\end{align*}
\]
The termination hypotheses for \( \text{quicksort} \) are computed as
\[
\forall x : \text{list} \quad x \neq \text{empty} \rightarrow \Delta^2 \text{smaller(head(x) \ \text{tail}(x)))} \lor \text{TRUE},
\]
\[
\forall x : \text{list} \quad x \neq \text{empty} \rightarrow \Delta^2 \text{larger(head(x) \ \text{tail}(x)))} \lor \text{TRUE}.
\]

A.5. Mergesort

function \( \text{distribute.odd}(x : \text{list}) : \text{list} \) 
\[
\begin{align*}
\text{if } x = \text{empty} & \text{ then empty} \\
\text{if } x \neq \text{empty} \land \text{tail}(x) = \text{empty} & \text{ then } x \\
\text{if } x \neq \text{empty} \land \text{tail}(x) \neq \text{empty} & \text{ then add(head(x) \ \text{distribute.odd}(\text{tail}(\text{tail}(x))))}
\end{align*}
\]
computes a copy of \( x \) with all elements on even positions removed. \( \text{distribute.odd} \) is a \( \Delta^1 \)-bounded algorithm with \( \Delta^1 \)-difference algorithm.
function \(\Delta^1\)distribute.odd(\(x : \text{list}\)) : bool \;
\begin{align*}
&\text{if } x = \text{empty} \text{ then } \text{false} \\
&\text{if } x \neq \text{empty} \land \text{tail}(x) = \text{empty} \text{ then } \text{false} \\
&\text{if } x \neq \text{empty} \land \text{tail}(x) \neq \text{empty} \text{ then } \text{true},
\end{align*}

i.e. \(\Delta^1\)distribute.odd(\(x\)) returns \text{true} iff \(x\) has at least two elements.

function distribute.even(\(x : \text{list}\)) : list \;
\begin{align*}
&\text{if } x = \text{empty} \text{ then } \text{empty} \\
&\text{if } x \neq \text{empty} \land \text{tail}(x) = \text{empty} \text{ then } \text{empty} \\
&\text{if } x \neq \text{empty} \land \text{tail}(x) \neq \text{empty} \text{ then} \\
&\quad \text{add(head(tail(x)) distribute.even(tail(tail(x))))}
\end{align*}

computes a copy of \(x\) with all elements on odd positions removed. distribute.even is a 1-bounded algorithm with 1-difference algorithm

function \(\Delta^1\)distribute.even(\(x : \text{list}\)) : bool \;
\begin{align*}
&\text{if } x = \text{empty} \text{ then } \text{false} \\
&\text{if } x \neq \text{empty} \text{ then } \text{true},
\end{align*}

i.e. \(\Delta^1\)distribute.even(\(x\)) returns \text{true} iff \(x\) has at least one element.

function mergesort(\(x : \text{list}\)) : list \;
\begin{align*}
&\text{if } x = \text{empty} \text{ then } \text{empty} \\
&\text{if } x \neq \text{empty} \land \text{tail}(x) = \text{empty} \text{ then } x \\
&\text{if } x \neq \text{empty} \land \text{tail}(x) \neq \text{empty} \text{ then} \\
&\quad \text{merge(mergesort(distribute.odd(x))} \\
&\quad \quad \text{mergesort(distribute.even(x))}).
\end{align*}

The termination hypotheses for mergesort are computed as
\begin{align*}
\forall x : \text{list} & \quad x \neq \text{empty} \land \text{tail}(x) \neq \text{empty} \rightarrow \Delta^1\text{distribute.odd}(x), \\
\forall x : \text{list} & \quad x \neq \text{empty} \land \text{tail}(x) \neq \text{empty} \rightarrow \Delta^1\text{distribute.even}(x).
\end{align*}

A.6. Heapsort

structure tip node(left : tree key : number right : tree) : tree,

function pop(h : tree) : tree \;
\begin{align*}
&\text{if } h = \text{tip} \text{ then } \text{tip} \\
&\text{if } h \neq \text{tip} \land \text{left}(h) = \text{tip} \text{ then } \text{tip} \\
&\text{if } h \neq \text{tip} \land \text{left}(h) \neq \text{tip} \land \text{depth(left}(h)) > \text{depth(right}(h)) \text{ then} \\
&\quad \text{node(pop(left(h)) key(h) right(h))}
\end{align*}
if \( h \neq \text{tip} \land \text{left}(h) \neq \text{tip} \land \text{depth}(\text{left}(h)) \leq \text{depth}(\text{right}(h)) \) then
\[ \text{node}(\text{left}(h) \text{ key}(h) \text{ pop}(\text{right}(h))) \]

computes a copy of \( h \) with the bottom node removed (if \( h \) is a heap). \text{pop} is a 1-bounded algorithm with 1-difference algorithm (after recursion elimination)

\[ \text{function } \Delta^1\text{pop}(h : \text{tree}) : \text{bool} \leftarrow \]
\[ \text{if } h = \text{tip} \text{ then false} \]
\[ \text{if } h \neq \text{tip} \text{ then true} . \]

\[ \text{function } \text{swap}(h : \text{tree}) : \text{tree} \leftarrow \]
\[ \text{if } h = \text{tip} \text{ then } \text{tip} \]
\[ \text{if } h \neq \text{tip} \land \text{left}(h) = \text{tip} \text{ then } \text{tip} \]
\[ \text{if } h \neq \text{tip} \land \text{left}(h) \neq \text{tip} \land \text{depth}(\text{left}(h)) > \text{depth}(\text{right}(h)) \text{ then} \]
\[ \text{node}(\text{pop}(\text{left}(h)) \text{ bottom}(h) \text{ right}(h)) \]
\[ \text{if } h \neq \text{tip} \land \text{left}(h) \neq \text{tip} \land \text{depth}(\text{left}(h)) \leq \text{depth}(\text{right}(h)) \text{ then} \]
\[ \text{node}(\text{left}(h) \text{ bottom}(h) \text{ pop}(\text{right}(h))) \]

replaces the key in the root of \( h \) by the bottom key of \( h \) and removes the bottom node from \( h \) (if \( h \) is a heap). \text{swap} is a 1-bounded algorithm with 1-difference algorithm

\[ \text{function } \Delta^1\text{swap}(h : \text{tree}) : \text{bool} \leftarrow \]
\[ \text{if } h = \text{tip} \text{ then false} \]
\[ \text{if } h \neq \text{tip} \land \text{left}(h) = \text{tip} \text{ then true} \]
\[ \text{if } h \neq \text{tip} \land \text{left}(h) \neq \text{tip} \land \]
\[ \text{depth}(\text{left}(h)) > \text{depth}(\text{right}(h)) \text{ then} \]
\[ \Delta^1\text{pop}(\text{left}(h)) \]
\[ \text{if } h \neq \text{tip} \land \text{left}(h) \neq \text{tip} \land \]
\[ \text{depth}(\text{left}(h)) \leq \text{depth}(\text{right}(h)) \text{ then} \]
\[ \Delta^1\text{pop}(\text{right}(h)) , \]

i.e. \( \Delta^1\text{swap}(h) \) returns \text{true} iff \( h \) has at least one node.

\[ \text{function } \text{sift}(h : \text{tree}) : \text{tree} \leftarrow \]
\[ \text{if } h = \text{tip} \text{ then tip} \]
\[ \text{if } h \neq \text{tip} \land \text{left}(h) = \text{tip} \text{ then } h \]
\[ \text{if } h \neq \text{tip} \land \text{left}(h) \neq \text{tip} \land \]
\[ \text{key}(h) = \text{min.key}(h \text{ left}(h) \text{ right}(h)) \text{ then } h \]
\[ \text{if } h \neq \text{tip} \land \text{left}(h) \neq \text{tip} \land \]
\[ \text{key}(\text{left}(h)) = \text{min.key}(h \text{ left}(h) \text{ right}(h)) \text{ then} \]
computes a heap consisting of all keys in h provided \text{left}(h) and \text{right}(h) are heaps and h has heap structure. \text{sift} is a 1-bounded algorithm with 1-difference algorithm (obtained after recursion elimination)

\begin{verbatim}
function \Delta^1\text{sift}(h : tree) : bool \Leftarrow
    if \text{true} then \text{false}.

function \text{heapsort}(h : tree) : list \Leftarrow
    if h = \text{tip} then \text{empty}
    if h \neq \text{tip} then \text{add(key(h) heapsort(sift(swap(h)))})
\end{verbatim}

computes a sorted list consisting of all keys in h if h is a heap. The termination hypothesis for \text{heapsort} is computed as

\[ \forall h : \text{tree} \quad h \neq \text{tip} \rightarrow \Delta^1\text{sift}(\text{swap}(h)) \lor \Delta^1\text{swap}(h). \]

Appendix B

We summarize the development of the estimation calculus presented in Section 4 with the following definition:

**Definition B.1 (Estimation calculus).** For an admissible specification S and a family \( \Gamma(S) \) of argument-bounded functions in S, the estimation calculus is given as:

1. **Language:** Estimation formulas, i.e. expressions for the form \( \langle t \leq_r \Delta \rangle \) such that \( t \) and \( r \) are terms from \( \mathcal{E}(\Sigma, \mathcal{V}) \), and \( \Delta \) is a quantifier-free formula.

2. **Axioms and inference rules** (estimation rules):
   - **identity:**
     \[
     \begin{array}{c}
     \hline
     \langle t \leq_r t, \text{false} \rangle \\
     \hline
     \end{array}
     \quad \text{for all } t ;
     \]
   - **equivalence:** for all irreflexive constructors \( \text{ircons}_1 \) and \( \text{ircons}_2 
     \[
     \begin{array}{c}
     \hline
     \langle \text{ircons}_1(t_1 \ldots t_n) \leq_r \text{ircons}_2(r_1 \ldots r_m), \text{false} \rangle \\
     \text{for all } t_i, r_j ;
     \hline
     \end{array}
     \]
(c) **strong estimation:** for all irreflexive constructors \( \text{ircons} \) and all reflexive constructors \( \text{rcons} \)

\[
\left< \text{ircons}(t_1 \ldots t_n) \leq_\sigma \text{rcons}(r_1 \ldots r_m), \text{TRUE} \right>
\]

for all \( t_i, r_j \);

(d) **strong embedding:** for all reflexive constructors \( \text{rcons} \)

\[
\left< t \leq_\sigma r_k, \Delta \right>
\]

\[
\left< \text{rcons}(r_1 \ldots r_k \ldots r_m), \text{TRUE} \right>
\]

for all \( t, r_j, \Delta \);

(e) **argument estimation:** for all \( f \in \Gamma_f(S) \)

\[
\left< t_p \leq_\sigma r, \Delta \right>
\]

\[
\left< f(t_1 \ldots t_n) \leq_\sigma r, \Delta \right>
\]

\[
\left< \text{rcons}(r_1 \ldots r_n), \text{TRUE} \right>
\]

for all \( t_i, r, \Delta \);

(f) **weak embedding:** for all reflexive constructors \( \text{rcons} \) with reflexive argument positions \( j_1, \ldots, j_h \)

\[
\left< t_{j_1} \leq_\sigma r_{j_1}, \Delta_{j_1} \right>, \ldots, \left< t_{j_h} \leq_\sigma r_{j_h}, \Delta_{j_h} \right>
\]

\[
\left< \text{rcons}(t_1 \ldots t_n) \leq_\sigma \text{rcons}(r_1 \ldots r_n), \Delta_j \right>
\]

for all \( t_i, r_j, \Delta_i \);

(g) **minimum:** for all irreflexive constructors \( \text{ircons} \) and for all reflexive constructors \( \text{rcons} \) with selectors \( \text{sel}_{j_1}, \ldots, \text{sel}_{j_{h(f)}} \)

\[
\left< \text{ircons}(t_1 \ldots t_n) \leq_\sigma r, \right>
\]

\[
\left< \text{rcons}(\text{sel}_{j_1}(r) \ldots \text{sel}_{j_{h(f)}}(r)) \right>
\]

\[
\left< \text{rcons}(\text{sel}_{j_1}(r) \ldots \text{sel}_{j_{h(n)}}(r)) \right>
\]

\[
\left< \text{rcons}(\text{sel}_{j_1}(r) \ldots \text{sel}_{j_{h(n)}}(r)) \right>
\]

(3) A deduction of \( \left< t_n \leq_\sigma r_n, \Delta_n \right> \) in the E-calculus is a sequence of estimation formulas \( \left< t_1 \leq_\sigma r_1, \Delta_1 \right>, \ldots, \left< t_n \leq_\sigma r_n, \Delta_n \right> \) such that each estimation formula \( \left< t_k \leq_\sigma r_k, \Delta_k \right> \) in the sequence either is an axiom or can be inferred by one of the inference rules from some estimation formulas preceding \( \left< t_k \leq_\sigma r_k, \Delta_k \right> \) in the sequence.

**Theorem 5.1.** A normal algorithm function \( f(x_1 : s_1 \ldots x_n : s_n) : s \leq \ldots \) strongly terminates in an admissible specification \( S \) if there exists some non-empty set \( P \subseteq \{1, \ldots, n\} \) such that for each recursive call \( f(t_1 \ldots t_n) \) in a case "if \( \varphi \) then . . . " of \( f \)

(1) \( t_i \leq_{\Gamma(f)} x_i \) for all \( i \in P \), and

(2) \( \forall x_1 : s_1 \ldots x_n : s_n \) \( \varphi \rightarrow \bigvee_{i \in P} \Delta_{\Gamma(f)}(t_i, x_i) \) \( \in Th(S) \).

**Proof.** Let \( <_f \) be the binary relation defined on \( \mathcal{T}(\Sigma)^{s_1 \ldots s_n} \) as \( (r_1 \ldots r_n) <_f (q_1 \ldots q_n) \) if \( \Sigma_{i \in P} \#_{s_i}(r_i) <_N \Sigma_{i \in P} \#_{s_i}(q_i) \) (cf. Section 4.1). Obviously \( <_f \) is well-founded and we prove that \( f \) strongly terminates with respect to \( <_f \) (cf.
Definition 3.3). Since only normal algorithms are considered here we have to verify that for each \((q_1 \ldots q_n) \in \mathcal{F}(\Sigma)_{t_1 \ldots t_n}\) and for each recursive call \(f(t_1 \ldots t_n)\) in a case "if \(\varphi\) then ..." of \(f\)
\[
\mathcal{J}[x_1 \ldots x_n/q_1 \ldots q_n] = \varphi \text{ implies } \\
\mathcal{J}[x_1 \ldots x_n/q_1 \ldots q_n](t_1 \ldots t_n) <_{f} (q_1 \ldots q_n), \quad (B.1)
\]
where \(\mathcal{J}\) is the standard model of \(S\). Assume that \((q_1 \ldots q_n) E \mathcal{F}(\Sigma)_{t_1 \ldots t_n}\) such that \(\mathcal{J}[x_1 \ldots x_n/q_1 \ldots q_n] = \varphi\). Then we conclude from requirement (1) and Theorem 4.4 that
\[
\#_{s_i}(\mathcal{J}[x_1 \ldots x_n/q_1 \ldots q_n](t_i)) \leq_{\mathcal{N}} \#_{s_i}(q_i) \quad \text{for each } i \in P. \quad (B.2)
\]
We obtain \(\mathcal{J}[x_1 \ldots x_n/q_1 \ldots q_n] = \bigvee_{i \in P} \Delta_{g} (t_i, x_i)\) by requirement (2), therefore \(\mathcal{J}[x_1 \ldots x_n/q_1 \ldots q_n] = \Delta_{g} (t_j, x_j)\) for some \(j \in P\), and with Theorem 4.4
\[
\#_{s_j}(\mathcal{J}[x_1 \ldots x_n/q_1 \ldots q_n](t_j)) <_{\mathcal{N}} \#_{s_j}(q_j) \quad \text{for some } j \in P \quad (B.3)
\]
must hold. We infer from \((B.2)\) and \((B.3)\)
\[
\sum_{i \in P} \#_{s_i}(\mathcal{J}[x_1 \ldots x_n/q_1 \ldots q_n](t_i)) <_{\mathcal{N}} \sum_{i \in P} \#_{s_i}(q_i) \quad (B.4)
\]
and \(\mathcal{J}[x_1 \ldots x_n/q_1 \ldots q_n](t_1 \ldots t_n) <_{f} (q_1 \ldots q_n)\) is verified. \(\square\)

**Theorem 6.2.** If \(S\) is an admissible specification containing a \(p\)-bounded algorithm function \(g(x_1 : s_1 \ldots x_n : s_n) : s \leftarrow \cdots\) and its \(p\)-difference algorithm function \(\Delta^p g(x_1 : s_1 \ldots x_n : s_n) : bool \leftarrow \cdots,\) then \(\text{TRUE}(g(x_1 \ldots x_n) \leq_{p} x_p, \Delta^p g(x_1 \ldots x_n))\).

**Proof.** The proof is by induction according to the recursions in the \(p\)-bounded algorithm. We find for each non-recursive case "if \(\varphi_i\) then \(r_i\)" of \(g\) that \(\models_{\mathcal{F}(S)} \langle [r_i] \leq_{p} [x_p], \gamma_i \rangle\), hence by Theorem 4.4, Lemma 5.6 and the definitions of \(g\) and \(\Delta^p g\)
\[
[\forall x_1 : s_1 \ldots x_n : s_n. \varphi_i \rightarrow g(x_1 \ldots x_n) \leq_{p} x_p] \in \text{Th}(S), \quad (B.5)
\]
\[
[\forall x_1 : s_1 \ldots x_n : s_n. \varphi_i \rightarrow \Delta^p g(x_1 \ldots x_n) \leftrightarrow g(x_1 \ldots x_n) <_{p} x_p] \in \text{Th}(S). \quad (B.6)
\]
For each recursive case "if \(\varphi_i\) then \(r_i\)" and for each recursive call \(g(t_1 \ldots t_n)\) in \(r_i\) we use
\[
\text{TRUE}(g([t_1] \ldots [t_n]) \leq_{p} [t_p], \Delta^p g(t_1 \ldots t_n)) \quad (B.7)
\]
as an induction hypothesis. For each term \(t\) and each formula \(\Delta\) we find that \((B.7)\) and \(\text{TRUE}(\langle [t_p] \leq_{p} t, \Delta \rangle\) implies
\[
\text{TRUE}(g([t_1] \ldots [t_n]) \leq_{p} t, \Delta^p g(t_1 \ldots t_n) \land \Delta). \quad (B.8)
\]
Hence we have proved that
is a sound estimation rule (relative to the recursive case under consideration). Consequently \( \vdash_{F(S)+} \langle [r_1] \leq_{\tau} [x_p], \gamma \rangle \) implies \( \text{TRUE} \langle [r_1] \leq_{\tau} [x_p], \gamma \rangle \) (cf. Theorem 4.4), and we conclude with Lemma 5.6 that (B.5) and (B.6) also hold for the recursive cases of \( g \).

Using the case-completeness of \( g \), i.e. \( \forall x_1 : s_1 \ldots x_n : s_n \varphi_1 \vee \cdots \vee \varphi_k \in \text{Th}(S) \), the statement of the theorem is obtained from (B.5) and (B.6). \( \square \)

References


