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A Predicate Transformer for Unification

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Abstract

In this paper we study unification as predicate transformer. Given a unification problem expressed as a set of sets of terms $U$ and a predicate $P$, we are interested in the strongest predicate $R$ (w.r.t. the implication) s.t. if $P$ holds before the unification of $U$ then $R$ holds when the unification is performed. We introduce a Dijkstra-style calculus that given $P$ and $U$ computes $R$. We prove the soundness, completeness and termination of the calculus. The predicate language considered contains monotonic predicates together with some non-monotonic predicates like $\text{var}$, $\neg\text{ground}$, $\text{share}$ and $\neg\text{share}$. This allows to use the calculus for the static analysis of run-time properties of Prolog programs.

1 Introduction

The standard view of logic programming is declarative, i.e. a program describes some predicate or function without referring to the way it will be computed. Nevertheless computational aspects become fundamental for the study of run-time properties of Prolog programs, like the actual form of the arguments of a goal before and after its call. In Prolog unification is the main computational mechanism since it produces the value of the variables during the execution of a goal in a program. To study its effect on the values of variables we study unification by means of predicate transformers. The use of predicate transformers for semantic analysis has been studied in the setting of imperative programming: it
was advocated by Floyd [5] and by Dijkstra [3] for program verification. The use of predicate transformers in the framework of logic programming is new. Given a unification problem expressed by a set of sets of terms \( U \), we introduce the predicate transformer \( sp.U \) such that \( sp.U.P \) is semantically equivalent to the strongest predicate \( R \) (w.r.t. implication) s.t. if \( P \) holds before the unification of \( U \), then \( R \) holds when the unification is performed. We show that \( sp.U.P \) could be computed in one step if \( P \) were a monotonic predicate. Since our aim is to infer run-time properties of Prolog programs, then the predicate language considered contains also non-monotonic predicates like \( \text{var} \) or \( \text{share} \). For this reason a careful analysis of some intermediate steps of the unification process is necessary. This yields to a non-trivial system of syntactic rules to compute \( sp.U.P \). The soundness, completeness and termination of the system is proved. The calculus can be used to infer run-time properties of logic programs. In Cousot and Cousot's original paper on abstract interpretation of imperative programs [2] everything was couched in terms of predicate transformers. Predicate transformers were used to define deductive semantics. Deductive semantics was used to design approximate program analysis frameworks. To propose a similar approach for logic programs we need the correspondent of program point for a logic program. In [7] Nilsson introduced a scheme for inferring run-time properties of logic programs based on a semantic description of logic programs that uses the concept of program point. We will show that the predicate transformer \( sp \) can be easily cast in such a theory. The rest of the paper is organized as follows. The next section contains some preliminaries and introduces the predicate transformer \( sp.U \). Section 3 introduces the transformation rules to compute \( sp.U.P \). In section 4 the soundness, completeness and termination of the calculus are proved. In section 5 we illustrate the use of the calculus for defining a forward semantics of Prolog programs.

2 Unification as Predicate Transformer

The computational meaning of unification in Prolog relies on the concept of substitution. A substitution is a mapping from variables to terms such that \( \text{dom}(\theta) = \{ v \mid v \neq \theta \} \) is finite. The notion of unification can be given w.r.t. a set of sets of terms [4] or w.r.t. a set of equations [6]. We choose the first approach. Let \( U \) be a finite set of sets of terms. A unifier for \( U \) is a substitution \( \theta \) such that every set in \( U \), under the application of \( \theta \), becomes a singleton, i.e. \( \forall S \in U \forall t, t' \in S \ (t \theta = t' \theta) \). A most general unifier for \( U \) is a unifier \( \bar{\theta} \) such that for every unifier \( \sigma \) there exists a substitution \( \gamma \) such that \( \bar{\theta} \gamma = \sigma \). The set of idempotent most general unifiers for \( U \) will be denoted by \( \text{mgu}(U) \). The operational meaning of \( U \) can be described as the partial function \( \lambda \alpha.\alpha.\mu \), where \( \alpha \) is a substitution and \( \mu \) is a fixed mgu in \( \text{mgu}(\alpha) \); clearly \( \lambda \alpha.\alpha.\mu \) is undefined if \( \text{mgu}(\alpha) = \emptyset \). We study unification by means of the predicate transformer \( sp.U \) (where \( sp \) stands for strongest postcondition [5]) with the following operational
meaning.

Definition 2.1 \( sp.U.P \) is true in precisely those substitutions \( \alpha \mu \) such that \( P\alpha \) is true and \( \mu \in \text{mgu}(U\alpha) \).

The choice to represent the unification process as set of sets of terms is motivated by the following observations:

\[
\text{mgu}\left(\{f(t_1, \ldots, t_n), f(s_1, \ldots, s_n)\}\right) = \text{mgu}\left(\{t_1, \ldots, t_n\}\right) \quad \text{and} \quad \text{mgu}(\{S_1, \ldots, S_n\}) = \text{mgu}(\{S_1 \cup S_2, S_3, \ldots, S_n\}) \text{ if } S_1 \cap S_2 \neq \emptyset.
\]

These two equalities will be used in our calculus for \( sp.U \) and they clearly lead to consider sets of sets of terms. For sake of clarity, we use double square brackets to enclose sets of terms \( S = [t_1, \ldots, t_m] \) and braces to enclose sets of sets of terms \( U = \{S_1, \ldots, S_n\} \).

We call a predicate \( P \) monotonic if it is (semantically) invariant under instantiation, that is for all substitutions \( \alpha, \beta \) if \( P\alpha \) is true then \( P\beta \) is true. Now let \( U \) be \( \{[t_1, \ldots, t_m], \ldots, [t^{m_k}_1, \ldots, t^{m_k}_m]\} \): we denote by \( U \) the predicate \( ((t_1^1 = \ldots = t_{m_1}^1) \land \ldots \land (t_{m_k}^{m_k} = \ldots = t_{m_k}^{m_k})) \). Then the following lemma holds.

Lemma 2.2 Let \( P \) be a monotonic predicate. Then \( P \land U \) is equivalent to \( sp.U.P \).

Proof. Let \( \alpha \) be s.t. \( P\alpha \) is true and let \( \mu \in \text{mgu}(U\alpha) \). Then \( U\alpha\mu \) is true and from \( P \) monotonic it follows that \( P\alpha\mu \) is true.

Viceversa let \( \alpha \) be s.t. \( (P \land U)\alpha \) is true. Then \( P\alpha \) is true and \( \epsilon \in \text{mgu}(U\alpha) \). So by Definition 2.1 \( (sp.U.P)\alpha \) is true.

Lemma 2.2 allows to compute \( sp.U.P \) when \( P \) is a monotonic predicate.

2.1 The Language

However we are interested also in properties that describe the structure of terms, like \( \text{var} \) or \( \neg \text{ground} \), since we want to use the predicate transformer to infer runtime properties of logic programs. Thus we introduce the language \( A \) defined on the alphabet containing the following classes of symbols:

- a countable set \( \text{VAR} \) of variables;
- a set \( \text{FUN} \) of functions;
- a set \( \text{PRED} = \text{Pred} \cup \{\text{free}, \text{var}, \neg \text{ground}, \text{share}, \neg \text{share}, \text{inst}\} \) of predicate symbols where \( \text{Pred} \) is a finite set of monotonic predicate symbols s.t. \( =, \neg \text{ground}, \neg \text{var}, \land, \neg \), \( \vdash \), \( \neg \text{invar} \) are in \( \text{Pred} \);
- the connectives \( \land \) and \( \lor \);
- the existential quantifier \( \exists \);
- the \text{connectives} ( and ) as punctuation symbols.

Variables will be normally denoted by the letters \( u, v, w, x, y, z \) (possibly subscripted or superscripted) and functions will be normally denoted by the
letters \( f, g, h \) (possibly subscripted). Let \( \text{TERM} \) be the set of terms built on \( \text{FUN} \) and \( \text{VAR} \). Terms will be normally denoted by the letters \( r, s, t \) (possibly subscripted or superscripted). Given a term \( t \), the set \( \text{vars}(t) \subseteq \text{VAR} \) denotes the set of variables that occur in \( t \). We call structured term a term of the form \( f(t_1, \ldots, t_m) \), where \( m \geq 1 \); we call proper subterm of \( t \) every subterm of \( t \) but \( t \). We assume that sequences are contained in \( \mathcal{A} \). We denote by \( t_1, \ldots, t_k \) and we write \( \langle t(k) \rangle \) or \( \langle t_1, \ldots, t_k \rangle \) if respectively the size or the elements of the sequence are relevant. Moreover we indicate with \( \xi \rho \) the sequence of terms obtained applying the substitution \( \rho \) to every element of the sequence \( \xi \).

We call atom a predicate of the form \( p(t_1, \ldots, t_n) \) where \( p \) is a predicate symbol of arity \( n \) and \( t_1, \ldots, t_n \) are terms. When ambiguity does not arise we write \( r(t_1, \ldots, t_n) \) as a shorthand for the predicate \( r(t_1 \land \ldots \land r(t_n)) \), where \( r \) is a predicate symbol of arity \( 1 \).

The truth value of a predicate \( P \in \mathcal{A} \) w.r.t. a substitution \( \alpha \) s.t. \( \text{vars}(P) \subseteq \text{dom}(\alpha) \) is defined inductively on the structure of \( P \), and the meaning of an atom is specified as follows:

- \( \lnot \text{var}(t) \alpha \) is true iff \( t \alpha \notin \text{VAR} \);
- \( \text{ground}(t) \alpha \) is true iff \( \text{vars}(t \alpha) = \emptyset \);
- \( \{ t_1 = t_2 \} \alpha \) is true iff \( t_1 \alpha = t_2 \alpha \) syntactically;
- \( \{ s \leq t \} \alpha \) is true iff \( s \alpha \) is a subterm of \( t \alpha \);
- \( \{ s < t \} \alpha \) is true iff \( s \alpha \) is a proper subterm of \( t \alpha \);
- \( \text{invar}(s, t) \alpha \) is true iff \( \text{vars}(s \alpha) \subseteq \text{vars}(t \alpha) \);
- \( \text{free}(s) \alpha \) is true iff \( s \alpha \in \text{VAR} \) and \( s \alpha \notin \text{vars}(y \alpha) \) for all \( y \in \text{dom}(\alpha) \) s.t. \( y \neq z \);
- \( \text{var}(z) \alpha \) is true iff \( z \alpha \in \text{VAR} \);
- \( \lnot \text{ground}(t) \alpha \) is true iff \( \text{vars}(t \alpha) \neq \emptyset \);
- \( \text{share}(s, t) \alpha \) is true iff \( \text{vars}(s \alpha) \cap \text{vars}(t \alpha) \neq \emptyset \);
- \( \lnot \text{share}(s, t) \alpha \) is true iff \( \text{vars}(s \alpha) \cap \text{vars}(t \alpha) = \emptyset \);
- \( \text{inst}(x, r_1, r_2, y) \alpha \) is true iff \( r_1 \alpha \) is the sequence \( \langle x_1, \ldots, x_m \rangle \), with \( x_i \in \text{var}(z \alpha) \) and \( x_i \notin \text{var}(y \alpha) \) for \( i \in [1, m] \), \( r_2 \alpha \) is the sequence \( \langle t_1, \ldots, t_m \rangle \) and \( \{ x_1 t_1, \ldots, x_m t_m \} \in \text{mgv}(\{ [\alpha, y \alpha] \}) \).

Notice that \( z \) and \( y \) in \( \text{inst}(x, r_1, r_2, y) \) represent two terms the second of which is an instance of the first. Thus the predicate \( \text{inst} \) expresses a special case of the unification.

Given two predicates \( P \) and \( Q \), we write \( P \equiv Q \) to indicate that \( P \) and \( Q \) are semantically equivalent. We can assume that the predicates \( \text{TRUE} \) (the predicate true w.r.t. all substitutions) and \( \text{FALSE} \) (the predicate false w.r.t. all substitutions) are in \( \mathcal{A} \), since \( \text{TRUE} \equiv (\text{var}(z) \lor \lnot \text{var}(z)) \) and \( \text{FALSE} \equiv (\text{var}(z) \land \lnot \text{var}(z)) \).

Predicates in \( \mathcal{A} \) are not in general monotonic, since all atoms built on predicate symbols not in \( \text{Pred} \) are non-monotonic by definition. So Lemma 2.2 is not sufficient to characterize \( \text{sp.U} \): consider for instance the unification \( \{ [x, a] \} \) and
the predicate var(x). Thus a careful analysis of the effect of the unification process on non-monotonic predicates is necessary. The fact that the connective ¬ is not in our language guarantees that atoms built on predicate symbols not in \( \mathit{Pred} \) are the only non-monotonic atoms of the language; this allows a case analysis of the effect of unification on non-monotonic predicates.

We introduce now some assumptions that will be used to simplify the form of the rules for \( \mathit{spM} \) that will be introduced in the next section. Predicates are of the form \( \exists x P \) where \( P \) doesn't contain any quantifier, it is in disjunctive normal form (i.e. it is a disjunction of conjunctions of atoms) and the equalities that occur in each conjunct are expressed by a set of equations in solved form. Atoms with predicate symbol \( \mathit{free}, \mathit{var}, \mathit{¬var}, \mathit{ground}, \mathit{¬ground}, \mathit{share}, \mathit{¬share}, \mathit{invar} \) have variables as arguments. For any formula \( \mathit{spM}.P \) the predicate \( P \) does not contain (existential) quantifiers.

All assumptions are not restrictive. Here the proof for the last one.

**Lemma 2.3** If the variable \( x \) does not occur in \( \mathit{U} \) then \( \mathit{spM}.\exists x P \) is equivalent to \( \exists x(\mathit{spM}.P) \) w.r.t. Definition 2.1.

**Proof.** Since \( x \) doesn't occur in \( \mathit{U} \) then the truth value of \( \exists x(\mathit{spM}.P) \) and \( (\mathit{spM}.\exists x P) \) does not depend on \( x \). Thus we can assume without loss of generality \( x \notin \mathit{dom}(\beta) \). Then \( (\mathit{spM}.\exists x P) \) is true iff there exist \( \alpha \) and \( \mu \) s.t. \( x \notin \mathit{dom}(\alpha), \mu \notin \mathit{dom}(\mu), \mu \in \mathit{mgu}(\mathit{U} \alpha), (\exists x P) \) is true and \( \beta = \alpha \mu \) iff there exist \( \alpha, \mu \) and \( t \) s.t. \( x \notin \mathit{dom}(\alpha), x \notin \mathit{dom}(\mu) \), \( \mu \in \mathit{mgu}(\mathit{U} \alpha) \), \( \mathit{P}(\alpha \cup \{x/t\}) \) is true and \( \beta = \alpha \mu \) iff there exist \( \alpha, \mu \) and \( t \) s.t. \( \mu \in \mathit{mgu}(\mathit{U} \alpha \cup \{x/t\}) \), \( \mathit{P}(\alpha \cup \{x/t\}) \) is true and \( (\beta \cup \{x/t\}) = (\alpha \cup \{x/t\}) \mu \) iff \( (\mathit{spM}.P)(\beta \cup \{x/t\}) \) is true. \( \square \)

**3 A Calculus for \( \mathit{spU} \)**

The following conditions on \( P \) and \( \mathit{U} \) characterize the types of formulas which will specify the scope of applicability of the rules for \( \mathit{spU}.P \).

(i) \( P \) is a conjunction of atoms.
(ii) For each equation \( x = t \) in \( P \), \( x \) does not occur in \( \mathit{U} \).
(iii) For every \( x \) occurring in \( \mathit{U} \) either \( \mathit{var}(x) \) or \( \mathit{¬var}(x) \) occurs in \( P \).
(iv) For all distinct variables \( x \) occurring in \( \mathit{U} \) and \( y \) occurring in \( P \) either \( \mathit{share}(x,y) \) or \( \mathit{¬share}(x,y) \) occurs in \( P \).
(v) \( \mathit{U} = \{S_1, \ldots, S_n\} \) contains disjoint sets, i.e. \( S_i \cap S_j = \emptyset \) for \( i \neq j \).
(vi) Each set in \( \mathit{U} \) contains more than one element.
(vii) Each set in \( \mathit{U} \) contains at most one structured element \( f(v_1, \ldots, v_m) \) and in such a case \( \mathit{free}(v_1), \ldots, \mathit{free}(v_m) \) occur in \( P \).
(viii) Every element \( x \) of a set \( S \in \mathit{U} \) is s.t. \( \mathit{free}(x) \) occurs in \( P \) if \( x \) occurs in the structured element of another set in \( \mathit{U} \) and \( \mathit{¬var}(x) \) occurs in \( P \) otherwise. Moreover, each set that contains a structured element also contains an element
ys.t. free(y) occurs in P. (Hence y occurs in the structured element of another set).

We introduce 3 types of formulas \( sp\mathcal{U}.P \) as follows.

- type 1: those which satisfy conditions (i)-(iii).
- type 2: those which satisfy conditions (i)-(vii).
- type 3: those which satisfy conditions (i)-(viii).

Each type of formula characterizes a simpler form of \( P \) and \( \mathcal{U} \). The final form will be a disjunction of formulas in the so called reduced form.

A formula \( sp\mathcal{U}.P \) is in reduced form if \( P \) is a conjunction of atoms, for each equation \( x = t \) in \( P \) does not occur in \( \mathcal{U} \), \( \mathcal{U} \) contains only disjoint sets of two or more variables, for all \( x \) occurring in \( \mathcal{U} \) both \( \neg\text{var}(x) \) and \( \neg\text{ground}(x) \) occur in \( P \) and for all \( x \) occurring in \( \mathcal{U} \) and \( y \) occurring in \( P \) either \( \text{share}(x, y) \) or \( \neg\text{share}(x, y) \) occurs in \( P \).

We are now ready to present the rules for \( sp\mathcal{U}.P \). The notation \( Ef \) will be used to indicate the formula obtained by replacing the occurrences of \( z \) in \( E \) with \( t \).

- If \( P = P_1 \lor ... \lor P_n \) then
  \[ sp\mathcal{U}.P \equiv sp\mathcal{U}.P_1 \lor ... \lor sp\mathcal{U}.P_n \]

- If \( x \) occurs in \( \mathcal{U} \) and neither \( \text{var}(x) \) nor \( \neg\text{var}(x) \) occurs in \( P \) then
  \[ sp\mathcal{U}.P \equiv sp\mathcal{U}.(P \land \text{var}(x)) \lor sp\mathcal{U}.(P \land \neg\text{var}(x)) \]

VAR1

- If \( P \) is a conjunction of atoms and \( z = t \) occurs in \( P \) then:
  \[ sp\mathcal{U}.P \equiv sp\mathcal{U}^x.P \]

EQ

- \( sp\mathcal{U}.\text{FALSE} \equiv \text{FALSE} \)

F

The following eight rules may be applied only to type 1 formulas.

- If \( x \) occurs in \( \mathcal{U} \) and \( y \) occurs in \( P \) and neither \( \text{share}(x, y) \) nor \( \neg\text{share}(x, y) \) occurs in \( P \) then
  \[ sp\mathcal{U}.P \equiv sp\mathcal{U}.(P \land \text{share}(x, y)) \lor sp\mathcal{U}.(P \land \neg\text{share}(x, y)) \]

SH1

- If \( U = \{f_1(s), f_2(t), \ldots, S_2, \ldots, S_n\} \) and \( f_1 \neq f_2 \), then
  \[ sp\mathcal{U}.P \equiv \text{FALSE} \]

MIS1

- If \( U = \{\{x_i, t_i\}, S_2, \ldots, S_n\} \) and either \( x \in \text{vars}(s) \) or the conjunct \( z < s \) occurs in \( P \) then
  \[ sp\mathcal{U}.P \equiv \text{FALSE} \]

MIS2
- If $U = \{[f(t_{i(k)}), \ldots, f(t_{m(k)})], S_2, \ldots, S_n\}$ then
  \[ sp.U.P \equiv sp.U'.P \]
  where $U' = \{[t_{i(k)}], S_2, \ldots, S_n\}$

- If $U = \{[f(t_{i(k)}), \ldots, f(t_{i(k)}), x_{i+1}, \ldots, x_m], S_2, \ldots, S_n\}$ with $i < m$ and either $i \geq 2$ or at least one $t_{i(k)}$ is not a variable or at least one $x_j$ is a variable such that $\neg \text{var}(x_j)$ occurs in $P$, then
  \[ sp.U.P \equiv \exists y_{i(k)}(sp.U'.P) \]
  where $U' = \{[f(y_{i(k)}), x_{i+1}, \ldots, x_m], [y_j, t_{j(k)}]_{j \in [i,k]}, S_2, \ldots, S_n\}$, $P' = P \land \text{free}(y_{i(k)})$ and $y_{i(k)}$ are fresh variables.

- If $U = \{[t, t_{i(m)}], [t, t_{i(m')}], S_2, \ldots, S_n\}$ then
  \[ sp.U.P \equiv sp.U'.P \]
  where $U' = \{[t, t_{i(m)}], [t, t_{i(m')}], S_2, \ldots, S_n\}$

- If $U = \{[t, S_2, \ldots, S_n]\} then
  \[ sp.U.P \equiv sp.U'.P \]
  where $U' = \{S_2, \ldots, S_n\}$

The following two rules may be applied only to type 2 formulas.

- If $U = \{[t, t_{i(m)}], S_2, \ldots, S_n\}$ where $x_m$ does not occur in the structured term of any set of $U$, $\text{var}(x_m)$ and $\neg \text{share}(x_m, y)$ occurs in $P$ for all $y \in \text{vars}(t)$, then
  \[ sp.U.P \equiv \exists z' \text{ sp}.U'.R \]
  where $U' = \{[t, \overline{z_{(m-1)}}, S_2, \ldots, S_n], R = (\bigwedge_{z \in \text{inst}(t), \text{var}(z_m)}z_m) \land P' \land z_m = t, z = (z \in \text{vars}(t) \mid P \Rightarrow \text{share}(z, z_m)), z' = z \rho$ is a variant of $z$ disjoint from $P$ and $P' = P^z_{x_m}$.

- If $U = \{[f(t_{i(k)}), x_{(m)}], S_2, \ldots, S_n\}$ and $\neg \text{var}(x_1), \ldots, \neg \text{var}(x_m)$ occur in $P$ then
  \[ sp.U.P \equiv \neg \exists y_{i(k)} sp.U'(P \land x_1 = f(y_{i(k)}^1) \land \ldots \land x_m = f(y_{i(k)}^m)) \]
  where $U' = \{[s_1, y_{i(k)}^m]_{i \in [i,k]}, S_2, \ldots, S_n\}$ and $y$ is the sequence $y_{i(k)}^1, \ldots, y_{i(k)}^m$ of fresh variables.

The following three rules may be applied only to type 3 formulas.
- If there is a set \( S \in U \) that contains a structured term then
\[
sp.U.P \equiv FALSE
\]  
\text{MIS3}

- If \( x \) occurs in \( U \) and neither \( \text{ground}(x) \) nor \( \neg \text{ground}(x) \) occurs in \( P \) then
\[
sp.U.P \equiv sp.U.(P \land \text{ground}(x)) \lor sp.U.(P \land \neg \text{ground}(x))
\]  
\text{GR1}

- If \( U = \{[x(m)], S_1, \ldots, S_n\} \) and \( \text{ground}(x_m) \) occurs in \( P \) then
\[
sp.U.P \equiv \exists y, \exists z, \exists \sigma \ sp.U'.R
\]  
\text{GR2}

where \( U' = \{S_2, \ldots, S_n\} \),
\[
\bar{x} = (x \in \text{vars}(P) | P \Rightarrow \text{share}(x, z_i) \text{ for some } i \in [1, n-1]), \bar{x'} = \bar{x}\rho \text{ is a variant of } \bar{x} \text{ disjoint from } P, z_x \text{ and } y_x \text{ are the sequences of fresh variables } z_x \text{ and } y_x \text{ with } z \in \bar{x}, P' = P_{\bar{x}'} \text{ and } R \text{ is the predicate}
\[
\Lambda_{x \in \bar{x}} \text{inst}(x\rho, z_x, y_x, z) \land_{x \in \text{vars}(P)} \text{isvar}(z_x, z_x) \land_{y \in \text{vars}(P)} \neg \text{share}(z_x, y)
\]
\[
\land P' \land z_1 = \ldots = z_m.
\]

To a formula in reduced form we can apply the following rule.

- If \( sp.U.P \) is in reduced form, where \( U = \{[x_1], \ldots, [x_n]\} \), then
\[
sp.U.P \equiv \exists y, \exists z, \exists \sigma \ (R \land U)
\]  
\text{RF}

where \( U \) is the predicate \((x_1 = \ldots = x_{m_1}) \land \ldots \land (x_n = \ldots = x_{m_n})\),
\[
\bar{x} = (x \in \text{vars}(P) | P \Rightarrow \text{share}(x, z_i) \text{ for some } i \in [1, m_i], j \in [1, n]) \}, \bar{x'} = \bar{x}\rho \text{ is a variant of } \bar{x} \text{ disjoint from } P, z_x \text{ and } y_x \text{ are the sequences of fresh variables } z_x \text{ and } y_x \text{ with } z \in \bar{x}, P' = P_{\bar{x}'} \text{ and } R = (\Lambda_{x \in \bar{x}} \text{inst}(x\rho, z_x, y_x, z) \land P').
\]

The previous rules are natural abstractions of the relative unification step except rules MIS3, VAR2, GR2 and RF. Rule MIS3 relies on the condition that the formula is of type 3 and \( U \) contains at least a set with a structured element. In this case it can be proven that \( U \) has no unifier.

Rules VAR2, GR2 and RF take into account how sharing among variables can propagate the bindings produced by the considered transformation and how the transformations affect the truth of the non-monotonic atoms. To keep track of the way the predicate is modified suitable variables are renamed with fresh variables existentially quantified and suitable predicates are introduced to specify the link among the original variables and the renamed ones.

All the rules are syntactic. Thus the set of rules provides a (nondeterministic) algorithm. We will see in the following section that this algorithm terminates and computes \( sp.U.P \). We conclude this section with some examples.

Let \( P = \text{free}(x, y) \) and \( U = \{[f(x), y], [g(y), z]\} \). Since \( sp.U.P \) is of type 3, then by rule MIS3 it is equivalent to \( FALSE \). In fact an occur check does occur.
Let $P = (\text{free}(x, y) \land \neg \text{share}(x, y))$ and $U = \{[y], [x]\}$. Since $sp.U.P$ is of type 2, then we can apply rule $\text{VAR2}$. We obtain

$$\exists x'(U.P, \langle x', \langle f(y)\rangle, z \rangle \land z = f(y)).$$

By rules $\text{SI}$ and $\text{RF}$ we obtain

$$\exists x'(U.P, x', \langle f(y)\rangle, z) \land z = f(y),$$

which is equivalent to $(\text{free}(y) \land x = f(y))$.

Let $P = (\text{ground}(y) \land \neg \text{var}(x) \land \neg \text{ground}(x) \land \neg \text{share}(x, y))$ and $U = \{[x, y]\}$. Since $sp.U.P$ is of type 3, then we can apply rule $\text{GR2}$. We obtain

$$\exists x', z_x y_x (U.P, x', \langle \text{var}(x, y) \rangle, z_x \land z = y).$$

By rule $\text{RF}$ we obtain

$$\exists x', z_x y_x U.P, x', \langle \text{var}(x, y) \rangle, z_x \land z = y,$$

which is equivalent to $(\text{ground}(y) \land \text{var}(y) \land y = z)$.

Let $P = (\neg \text{var}(x, y) \land \neg \text{ground}(x, y) \land \text{share}(x, y))$ and $U = \{[x, y]\}$. Since $sp.U.P$ is in reduced form, then we can apply rule $\text{RF}$. We obtain

$$\exists x', y_x z_x y_x U.P, x', \langle \text{var}(x, y) \rangle, z_x \land z = y,$$

which is equivalent to $(x = y \land \neg \text{var}(x, y))$, if $\text{CON}$ contains at least a function of arity greater than one and a constant; otherwise it is equivalent to $(x = y \land \neg \text{var}(x, y) \land \neg \text{ground}(x, y))$.

### 4 Soundness and Completeness of the Calculus

We indicate by $\mathcal{H}_p$ the set of rules but $\text{RF}$. We first show that all the rules are equivalences. Then we show that a formula $sp.U.P$ can be reduced in a finite number of steps to a disjunction of formulas in reduced form, by applying rules from $\mathcal{H}_p$. Finally rule $\text{RF}$ applied to each disjunct will give the desired predicate (of $A$) relative to $sp.U.P$.

**Theorem 4.1** All rules are equivalences (with respect to Definition 2.1)

**Proof.** The proof is not difficult except for rules $\text{MIS3}$, $\text{VAR2}$, $\text{GR2}$ and $\text{RF}$ which have a quite technical proof.

**MIS3** By hypothesis the formula is of type 3 and $U$ contains at least a set with a structured element. Then by condition (vii) each set that contains a structured element $f(y_1, \ldots, y_k)$ also contains at least $f$ variable $z$ that occurs in the structured element of another set. In such a situation we can eventually extract from $U$ a subset $\{S_1, \ldots, S_k\}$ of sets such that

- $S_1 = \{f_1([x], y_1, \ldots), z_1, \ldots\}$
- $S_2 = \{f_2([x], x_1, \ldots), z_2, \ldots\}$
... 

\[ S_t = [t_t(t_1, ..., t_{t-1}, ..., t_{t})]. \]

Clearly \( \{s_1, ..., s_t\}^a \) has no unifier.

In the next proofs we use the following properties of most general unifiers:

1) Let \( U = \{S_1, ..., S_n\} \). If \( \beta \in \text{mgu}(U) \) and \( \mu \in \text{mgu}(U\beta) \) then \( \mu \cup \beta \in \text{mgu}(U) \), where \( \mu = (\beta\mu)_{\text{dom}(\beta)} \).

2) \( \text{mgu}([S_1, ..., S_n]) = \text{mgu}(\{S_2, ..., S_n\}) \).

**VAR2** Let \( \alpha \) be such that \( Pa \) is true and let \( \mu \in \text{mgu}(U\alpha) \). Let \( \alpha' \) be such that

\[
\begin{align*}
za' &= \begin{cases}
  x_\alpha & \text{if } x = z_m, \\
  x_\alpha & \text{if } x = z_p, \\
  (x_\alpha)^{z_m -}\alpha' & \text{otherwise}.
\end{cases}
\end{align*}
\]

Let \( A' \) be an atom in \( P' \). Then \( A' = A' \) with \( A \) atom in \( P \). If \( A' \) is monotonic then \( A' \alpha' \) is an instance of \( A\alpha \). Otherwise \( A' \alpha' = A\alpha \). Thus in both cases \( A' \alpha' \) is true. From \( \alpha' = \alpha' \) it follows that \( \text{inst}(z_p, (z_m)^{z_m -}\alpha', (t), z_\alpha') \) and \( (z_m) = t) \alpha' \) are both true. Then \( R\alpha' \) is true. Now let \( \mu' \) be s.t. \( \mu = \mu' \cup \{z_m \alpha / z_\alpha' \} \). Then from \( U\alpha' = (U\alpha)^{z_m -}\alpha' \) it follows by property 1) that \( \mu' \in \text{mgu}(U\alpha') \). Thus \( (sp\{U\alpha' \} \alpha') \alpha' \) is true and, since \( za' = z_\alpha' \mu' \) for all \( z \) in \( P \), then \( (3\alpha' \alpha') \alpha' \) is true.

Viceversa let \( \alpha' \) be such that \( R\alpha' \) is true and let \( \mu' \in \text{mgu}(U\alpha') \). Let \( \alpha \) be such that

\[
\begin{align*}
za = \begin{cases}
  z_\alpha' & \text{if } z \in z_1, \\
  z_\alpha' & \text{otherwise}.
\end{cases}
\end{align*}
\]

Then \( Pa = P\alpha' \) is true. Let \( \mu = \mu' \cup \{z_m \alpha / z_\alpha' \} \). By \( \text{inst}(z_p, (z_m)^{z_m -}\alpha', (t), z_\alpha') \) true for all \( z \) it follows that \( U\alpha' = (U\alpha)^{z_m -}\alpha' \). Then by property 1) \( \mu \in \text{mgu}(U\alpha) \). Thus \( (sp\{U\alpha' \} \alpha') \alpha' \) is true and, since \( z_\alpha' \mu' = (t), z_\alpha' \) and \( (z_m) = t) \alpha' \) are both true. Then \( (3\alpha' \alpha') \alpha' \) is true.

**GR2** Let \( \alpha \) and \( \mu \) be such that \( Pa \) is true and \( \mu \in \text{mgu}(U\alpha) \), let \( \mu_i = \mu_{i+1} \). From \( \text{ground}(z_\alpha) \) true it follows that \( z_\alpha \mu = z_\alpha \mu_i \) for \( i \in [1, m - 1] \). Let \( \alpha' \) be s.t.

\[
\begin{align*}
\omega_\alpha' &= \begin{cases}
  z_\alpha & \text{if } w = z_p \text{ with } z \in z_1, \\
  z_\alpha & \text{if } w = z, \text{ for } i \in [1, m_1], \\
  (y_1 \cdots y_{m_1} - \mu) & \text{if } w = z_\alpha \text{ with } z \in z_1, \\
  (\omega_\alpha) \mu_1 \cdots \mu_{m_1} & \text{otherwise}.
\end{cases}
\end{align*}
\]

where \( y_i \) is the sequence of variables in \( \text{vars}(z_\alpha) \cap \text{vars}(z_\alpha) \) for \( i \in [1, m - 1] \). Let \( A' \) be an atom of \( P' \). Then \( A' = A' \) with \( A \) atom in \( P \). If \( A' \) is monotonic then \( A' \alpha' \) is an instance of \( A\alpha \). If \( A' \) is non-monotonic then \( A' \alpha' = A\alpha \). In both cases \( A' \alpha' \) is true. Moreover \( (z_1 = \cdots = z_m) \alpha' \) is true because \( z_m \alpha' = z_m \alpha = z_\alpha \) for all \( i \in [1, m - 1] \), which is true because all variables in \( z_\alpha \) occur in \( z_\alpha \) for some \( i \in [1, m - 1] \) and \( \alpha' \alpha' \) is obtained replacing the variables in...
all $x\alpha$ with ground terms, \(\text{inst}(zp, z_x, y_x, z)\alpha'\) is true because $z\alpha' = z\alpha$, $z\alpha' = (z\alpha )z'z$, and $\neg \text{share}(z_x, z)\alpha'$ true imply \(\{z\alpha'/y_x\alpha'\} \in \text{mgu}({\{z\alpha', z\alpha\}})\); finally \(\text{invar}(z_x, \{z_x, \ldots, z_{m-1}\})\alpha'\) is true by construction. Thus $Ra'$ is true. Now let $\mu' = \mu|_{\text{vars}(U\alpha')}$. We have that $\mu_1 \ldots \mu_{m-1}$ is in $\text{mgu}({\{x_1, \ldots, z_m\}}\alpha)$, \(\text{range}(\mu_1 \ldots \mu_{m-1}) = \emptyset\) because $m\alpha$ is ground, $U\alpha' = U\alpha'\mu_1 \ldots \mu_{m-1}$. Then $\mu = \mu' \cup \mu_1 \ldots \mu_{m-1}$ and by properties 1) and 2) $\mu'$ is in $\text{mgu}(U\alpha')$. Then $(spU'.R)\alpha'\mu'$ is true and, since $z\alpha'\mu' = z\alpha\mu$ for all $z$ occurring in $P$, then $(\exists z_x, z_x, y_x spU'.R)\alpha\mu$ is true.

Viceversa let $\alpha'$ be such that $Ra'$ is true and let $\mu' \in \text{mgu}(U\alpha')$. Let $\alpha$ be s.t.

\[ z\alpha = \left\{ \begin{array}{ll} z\alpha' & \text{if } z \in z, \\ z\alpha' & \text{otherwise.} \end{array} \right. \]

Then $P\alpha = P'\alpha'$ is true. Let $\mu = \mu' \cup \beta$ with $\beta = \{\{x_\alpha, \alpha'/y_x\alpha'\}_{i \in [1, m-1]}\}$. From \(\text{inst}(zp, z_x, y_x, z)\alpha'\), $\neg \text{share}(z_x, z)\alpha'$ and $\text{invar}(z_x, \{z_x, \ldots, z_{m-1}\})\alpha'$ true it follows that $z\alpha' = z\beta$ for all $x \in x$. If $z \notin z$ then from $\neg \text{share}(z_x, z)\alpha'$ true for all $y$ it follows that $z\alpha' = (z\alpha')\beta = z\beta$. Then $z\alpha' = z\beta$ for all $x$ occurring in $P$. Then $U\alpha' = U\alpha\beta$. From $z_\alpha' = \ldots = z_{m}\alpha'$ true, $z_m\alpha'$ ground and \(\text{inst}(z_\alpha, z_{1}, y_x, z_{1})\alpha'$ true for all $i \in [1, m-1]$ it follows that $\beta \in \text{mgu}({\{x_1, \ldots, z_m\}}\alpha)$. Then by properties 1) and 2) it follows that $\mu \in \text{mgu}(U\alpha)$. Thus $(spU'.P)\alpha\mu$ is true and, since $z\alpha'\mu' = (z\alpha)\beta\mu' = z\alpha\mu$ for all $x$ occurring in $P$, then $(spU.P)\alpha'\mu'$ is true.

**RF** Let $\alpha$ and $\mu$ be such that $P\alpha$ is true, $\mu \in \text{mgu}(U\alpha)$. Let $\alpha'$ be s.t.

\[ u\alpha' = \left\{ \begin{array}{ll} \alpha & \text{if } w = zp \text{ with } z \text{ in } z, \\ u\alpha & \text{if } w \text{ occurs in } P, \\ y & \text{if } w = z_x \text{ with } z \text{ in } z, \\ y\mu & \text{if } w = y_x \text{ with } z \text{ in } z. \end{array} \right. \]

where $y$ is the sequence of variables occurring in $\text{dom}(\mu|_{\text{vars}(z\alpha)})$. Now $U\alpha'$ is true because $z_\alpha z' = z_\alpha u\mu$ for every $i \in [1, m_1]$, $j \in [1, n]$. Let $A'$ be an atom of $P'$. Then $A' = A'_E$ with $A$ atom in $P$. If $A'$ is monotonic then $A'\alpha'$ is an instance of $A\alpha$. If $A'$ is non-monotonic then $A'\alpha'$ is $A\alpha$. In both cases $A'\alpha'$ is true. Moreover \(\text{inst}(zp, z_x, y_x, z)\alpha')\alpha'$ is true because $z\alpha' = z\alpha$, $z\alpha' = z\alpha\mu$ and the substitution relative to the two sequences $z_x\alpha'$ and $y_x\alpha'$ is equal to $\mu|_{\text{vars}(z\alpha)}$. Since $\mu$ is idempotent by hypothesis, then $\mu|_{\text{vars}(z\alpha)} \in \text{mgu}({\{z_\alpha, z_\alpha\mu\}})$. Then $(R \land U)\alpha'$ is true and, since $z\alpha' = z\alpha\mu$ for every $x$ occurring in $P$, then $(\exists z_x, z_x, y_x(R \land U))\alpha\mu$ is true.

Viceversa let $\alpha'$ be s.t. $(R \land U)\alpha'$ is true. Let $\alpha$ be s.t.

\[ z\alpha = \left\{ \begin{array}{ll} z\alpha' & \text{if } z \in z, \\ z\alpha' & \text{otherwise.} \end{array} \right. \]

Then $P\alpha = P'\alpha'$ is true. Let $\mu$ be the substitution relative to the sequences $z_\alpha\alpha', y_x\alpha'$ for all $i \in [1, m_1]$, $j \in [1, n]$. Then $\mu \in \text{mgu}(U\alpha)$. Thus $(spU.P)\alpha\mu$
is true and, since $x\alpha u = x\alpha' u = x\alpha'$ for every $x$ that occurs in $P$, then $(sp_u P)\alpha'$ is true.$\Box$

**Theorem 4.2** The system $\mathcal{H}_{sp}$ is terminating.

**Proof.** (Sketch)

We show that no proof tree built using $\mathcal{H}_{sp}$ has an infinite branch. Rules $F$, $M1S1$, $M1S2$ and $MIS3$ have a predicate as right hand side, so they cannot belong to an infinite branch. To prove that only finitely many applications of the remaining rules are allowed, consider the tuple $r = (\text{leq, comp, funct, elem, disj, unvar, unshare, unground})$ of natural numbers with the lexicographic order. A structured term $f(t_1, \ldots, t_n)$ will be called compound if either some $t_i$ is not a variable or the variables $t_1, \ldots, t_n$ are not distinct. Then $\text{leq}$ denotes the number of variables in $U$ that occur as left hand side of an equation in $P$, $\text{comp}$ denotes the number of occurrences of compound subterms of terms in $U$, $\text{funct}$ denotes the number of occurrences of functor symbols in $U$, $\text{elem}$ denotes the total number of elements in the sets of $U$, $\text{disj}$ denotes the number of disjuncts in the disjunctive normal form of $P$, $\text{unvar}$ denotes the number of variables $x$ in $P$ such that neither $P \Rightarrow \text{var}(x)$ nor $P \Rightarrow \neg \text{var}(x)$ holds, $\text{unshare}$ denotes the number of variables $x$ in $P$ such that neither $P \Rightarrow \text{share}(x, y)$ nor $P \Rightarrow \neg \text{share}(x, y)$ holds for some variable $y$ distinct by $x$, $\text{unground}$ denotes the number of variables $x$ in $P$ such that neither $P \Rightarrow \text{ground}(x)$ nor $P \Rightarrow \neg \text{ground}(x)$ holds.

It is not difficult to check that the application of every rule of $\mathcal{H}_{sp}$ decreases the value of $r$. $\Box$

**Corollary 4.3** Rules of $\mathcal{H}_{sp}$ transform $sp_u P$ in a (semantically unique) disjunction of formulas in reduced form.

**Proof.** (Sketch)

By Theorem 4.1 all transformations are equivalences (w.r.t. Definition 2.1). By Theorem 4.2 there is a final form. Thus the final form is semantically unique. By contraposition it is not difficult to show that if the final form is not a disjunction of formulas in reduced form then one of the rules in $\mathcal{H}_{sp}$ may be applied. $\Box$

**5 Applications**

Predicate transformers are related to the core of abstract interpretation of imperative programs. In [2] predicate transformers are used to define deductive semantics. Deductive semantics is used to design approximate program analysis frameworks. To propose a similar approach in the setting of logic programming we need the correspondent of program point for a logic program. In [7] Nilsson introduced a scheme for inferring run-time properties of logic programs based on
a semantic description of logic programs that uses the concept of program point. The predicate transformer \( sp \) can be easily cast in such a theory. A clause of a logic program \( P \) is interpreted as a sequence of procedure calls. To each call \( A \) there corresponds a calling point \( \hat{x}A \) and a success point \( \hat{y}A \). The leftmost and rightmost points in the body of a clause \( C \) are called respectively entry- and exit points of the clause and are indicated respectively by \( \hat{x}C \) and \( \hat{y}C \). Goals are represented as elements of the set \( \text{Goals} := (\text{Points} \times \text{Env})^* \), where \( \text{Points} \) denotes the set of program points of \( P \) and \( \text{Env} \) is the set of predicates \( A \).

A transition system for \( P \) can be defined through two state transition schemes that transform elements of \( \text{Goals} \) as follows.

\[
(C; R) \xrightarrow{y} y, \\
(\hat{x}A; R) \xrightarrow{y} ((\hat{x}C; \text{TRUE}) :: (\hat{y}A; R) :: y)(T \sigma^{-1}),
\]

where \( A \) is a body atom, \( C \sigma \) is a variant of a clause \( C \) of \( P \) s.t. \( \text{vars}(\hat{x}A; R) :: y) \cap \text{vars}(C \sigma) = \emptyset \), \( T \equiv \text{sp}(\{\hat{x}A, \text{head}(C \sigma)\}) \cdot (R \land \text{free}(\text{vars}(C \sigma))) \neq \text{FALSE} \).

We assume that the program clauses are disjoint and that the definition of \( U \) in \( sp \) is generalized in the obvious way to atoms or terms. The application of a predicate \( R \) to a \( C \)-goal is defined as follows:

\[
(\text{nul})R = R, \\
((x; T) :: y)R = (x; T \bullet R) :: yR,
\]

where \( T \bullet R \) is (equivalent to) \( T' \land R \), with \( T' \) the strongest assertion (w.r.t. implication) s.t. \( T \rightarrow T' \) and \( (T' \land R) \neq \text{FALSE} \). Notice that \( T \bullet R \) is defined when \( R \) is consistent. For instance if \( T = (x = y \land \text{var}(x)) \) and \( R = \text{ground}(y) \) then \( T \bullet R \equiv (x = y \land \text{ground}(y)). \)

The previous transition schemes are obtained from those in [7] by taking as environment \( \text{Env} \) predicates instead of substitutions, by using the predicate transformer \( sp \) instead of the mgu as operation in the transition and the operation \( \bullet \) to model the application of a predicate to a \( C \)-goal.

To each program point \( i \) is associated a set \( \Theta_i \) of states which specifies when the program point becomes current. The set of states is defined as \( \text{Goals} \times \text{Goals} \), where the first component describes the \( C \)-goal that invoked the clause containing point \( i \) and the second component is the \( C \)-goal when the point is current. The semantics of \( P \) is defined as the least fixpoint of the system of equations relative to its program points. Every program point is either the entry point of a clause or the success point of a body atom. Then it is sufficient to define the meaning of entry- and success points:

\[
\Theta_C = \bigcup_{A \rightarrow C} \{(G_i; G_{i+1}) \mid \exists G((G_i; G_{i+1}) \in \Theta_A \land G_i \in C \land G_{i+1})\}, \\
\Theta_A = \bigcup_{A \rightarrow C} \{(G_i; \text{tail}(G_j)) \mid \exists G((G_i; G_j) \in \Theta_A \land (G_i; G_j) \in \Theta_C)\}.
\]

**Example** Consider the following simple case of concatenation of two lists:

\[
C_0 : \leftarrow \text{append}([a], [], z) \rightarrow \text{append}([a], [], z) \rightarrow \text{append}([a], [], z).
\]

\[
C_1 : \text{append}([H|L_1], L_2, [H|L_3]) \rightarrow \text{append}(L_1, L_2, L_3).
\]
Here the program points are explicitly labelled by integers. The meaning of this program, when append([a], [], z) is called with z free variable, can be given as least fixpoint of the following set of equations, where we use the notation of [7].

\[
\begin{align*}
\Theta_1 &= \{(\text{nil}; (\text{Co}; \text{free(z)}) :: \text{nil})\}, \\
\Theta_2 &= \{(G; \text{tail}(G_i)) \mid \exists G_i ((G; G_i) \in \Theta_1 \land ((G_i; G_j) \in \Theta_1 \lor (G_i; G_j) \in \Theta_4))\}, \\
\Theta_3 &= \{(G_i; G_{i+1}) \mid \exists G ((G; G_i) \in \Theta_1 \lor (G; G_i) \in \Theta_3) \land G_i \models C_i \land G_{i+1} \models C_{i+1})\}, \\
\Theta_4 &= \{(G; \text{tail}(G_i)) \mid \exists G_i ((G_i; G_j) \in \Theta_3 \land ((G_i; G_j) \in \Theta_4 \lor (G_i; G_j) \in \Theta_4))\}, \\
\Theta_5 &= \{(G_i; G_{i+1}) \mid \exists G ((G; G_i) \in \Theta_1 \lor (G; G_i) \in \Theta_3) \land G_i \models C_i \land G_{i+1} \models C_{i+1})\}.
\end{align*}
\]

Notice that in this case the fixpoint can be computed in finite time since the program terminates. We first calculate \(\Theta_3\). We need to compute

\[
\begin{align*}
\text{sp.}\{(\text{append}([a], [], z), \text{append}([H|L1], [L2, H|L3]), \text{free(z)}, H, L1, L2, L3)\}.
\end{align*}
\]

By rule STR1, rule VAR2 applied to \(L2\) and \(z\), rules SI, STR1 and rule VAR2 applied to \(H\) and \(L1\) we obtain the predicate

\[
T \equiv (H = a \land L1 = [] \land L2 = [] \land z = [a|L3] \land \text{free}(L3)).
\]

Since \((\text{free(z)} \cdot T) = T\) then \((\text{Co}; \text{free(z)}) :: \text{nil} \models C_i \land (\text{Co}; T) :: (\text{Co}; T) :: \text{nil}.

By rules STR1 and MIS1

\[
\begin{align*}
\text{sp.}\{(\text{append}([L1, L2, L3]), \text{append}([H'|L1'], [L2', [H'|L3']])\}.
\end{align*}
\]

is equivalent to FALSE. Hence

\[
\begin{align*}
\Theta_3 &= \{(\text{Co}; \text{free(z)}) :: \text{nil}, (\text{Co}; T) :: (\text{Co}; T) :: \text{nil}\}.
\end{align*}
\]

Consider now \(\Theta_5\). We need to compute

\[
\begin{align*}
\text{sp.}\{(\text{append}([], [L], \text{append}([L1, L2, L3]), \text{free(L)} \land T)\}.
\end{align*}
\]

By rule STR1, rule EQ applied to \(L1\) and \(L2\), rule SI, rule SH2 applied to \(L\) and rule VAR2 applied to \(L\) and \(L3\) we obtain the predicate

\[
R \equiv ((H = a \land L1 = [L2 = [L3 = [] \land z = [a]).
\]

Since \(T \cdot R = R\) then

\[
\begin{align*}
(\text{Co}; T) :: (\text{Co}; T) :: \text{nil} \models C_i \land (\text{Co}; R) :: (\text{Co}; R) :: \text{nil}.
\end{align*}
\]

By rules STR1 and MIS1

\[
\begin{align*}
\text{sp.}\{(\text{append}([a], [], z), \text{append}([], [L], L))\} \cdot \text{free}(L) \land T
\end{align*}
\]

is equivalent to FALSE. Hence

\[
\begin{align*}
\Theta_5 &= \{(\text{Co}; T) :: (\text{Co}; T) :: \text{nil}, (\text{Co}; T) :: (\text{Co}; T) :: (\text{Co}; T) :: (\text{Co}; T) :: \text{nil}\}.
\end{align*}
\]

Finally \(\Theta_2\) and \(\Theta_4\) can be easily calculated.
\[ \Theta_2 = \{ (\text{nil}; \langle C_{04}; R \rangle) \Rightarrow \text{nil}) \}; \]

\[ \Theta_4 = \{ (\langle \text{free}(z); [\text{nil}; \langle C_{14}; R \rangle \Rightarrow \langle C_{04}; R \rangle \Rightarrow \text{nil}) \}. \]

Every set \( \Theta_i \) describes the states associated to the program point \( i \). Thus for instance \( \Theta_3 \) specifies that the program point 3 becomes current only when the goal \( \text{append}([a], [\text{nil}], z) \) invokes \( C_1 \) with \( z \) free variable and in such a case \( H \) becomes equal to \( a \), \( L_1 \) and \( L_2 \) become equal to the empty list \([\text{nil}]\) and \( L_3 \) remains a free variable.

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