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HOL-$\lambda\sigma$: an intentional first-order expression of higher-order logic

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We give a first-order presentation of higher-order logic based on explicit substitutions. This presentation is intentionally equivalent to the usual presentation of higher-order logic based on λ-calculus, i.e. a proposition can be proved without the extensionality axioms in one theory if and only if it can be in the other. We show that the *Extended Narrowing and Resolution* first-order proof-search method can be applied to this theory. We get this way a step by step simulation of higher-order resolution. Hence expressing higher-order logic as a first-order theory and applying a first-order proof search method is a relevant alternative to a direct implementation. In particular, the well studied improvements of proof search for first-order logic could be reused at no cost for higher-order automated deduction. Moreover as we stay in a first-order setting, extensions, such as equational higher-order resolution, may be easier to handle.

Introduction

Higher-order logic is a formalism that allows a natural expression of program specifications and of mathematics. It is used in many theorem provers such as HOL, Isabelle, PVS, λ-Prolog, etc. In this paper, we are concerned with the automation of proof search in this logic.

Higher-order logic can be expressed in many different ways using combinators, λ-calculus, etc. Some of these formulations, but not all, present higher-order logic as a first-order theory. Such a formulation allows to use standard first-order methods for proof search and it may allow to handle extensions more easily. There are several ways to encode higher-order logic as a first-order theory and several proof search methods for each encoding, which are more or less efficient. For instance, higher-order logic can be encoded as a first-order theory using combinators. The theory obtained this way is equivalent to the standard presentation using λ-calculus, but it is not *intentionally* equivalent to it: some proofs use the extensionality axioms with combinators, while they do not in the standard presentation. Moreover, the additional use of the extensionality axioms which

express that two pointwise equivalent functions are equal and that two sets that have the same elements are equal, induces inefficiencies in proof search.

In this paper, which is a revised and extended version of (Dowek et al., 1999) and builds upon (Dowek et al., 1995) and (Dowek et al., 1998), we give a new first-order presentation of higher-order logic called HOL-$\lambda\sigma$. It uses the fact already noticed by several authors that explicit substitutions simplify algorithms and speed up implementations (Nadathur and Wilson, 1990; Magnusson, 1994; Dowek et al., 1995; Dowek et al., 1996; Muñoz, 1997b; Nadathur and Wilson, 1998). Making at work the *calculus of explicit substitutions* defined and studied in (Abadi et al., 1991; Curien et al., 1996), we first define HOL-$\lambda\sigma$ and show that it is intentionally equivalent to the usual presentation of higher-order logic based on λ-calculus, i.e. the theories are still equivalent when we drop the extensionality axioms in both cases.

We then show that proof-search in this theory can be mechanized with the *Extended Narrowing and Resolution* (ENAR) method introduced in (Dowek et al., 1998). The proof search method for higher-order logic obtained this way is as efficient as higher-order resolution and in fact simulates it step by step: proof search steps correspond and β-reduction steps correspond to $\lambda\sigma$-reduction steps. It keeps however the simplicity of first-order frameworks and could easily be extended, for instance with equational axioms.

At last, a rather surprising side effect of this presentation of higher-order logic is that it provides a clarification of the intricate skolemization rule of higher-order logic (Miller, 1983; Miller, 1987).

HOL-$\lambda\sigma$ and the ENAR proof search method rely upon a presentation of first-order logic called *deduction modulo* that allows to build-in a congruence identifying not only terms but also propositions. This leads to shorter and more direct proofs by identifying congruent propositions instead of requiring explicit equivalence proofs. Hence, we shall express HOL-$\lambda\sigma$ in deduction modulo. In order to remain self contained, we recall the principal ideas of deduction modulo in section 1. Then, we recall in section 2 the usual presentation of higher-order logic based on λ-calculus (HOL-λ) and in section 3 its first-order presentation based on combinators. Section 4 introduces HOL-$\lambda\sigma$ and establishes its main properties (termination, confluence, consistency and cut elimination). Section 5 is dedicated to the equivalence theorem between HOL-λ and HOL-$\lambda\sigma$ (which rests upon cut elimination). In section 6 we show that the rather intricate Skolem theorem for higher-order logic can be deduced from the first-order one. Finally, section 7 presents the ENAR proof search method (whose completeness rests also upon cut elimination) and its application to HOL-$\lambda\sigma$.

1. Deduction modulo

In this paper we shall use a presentation of first-order logic, called *deduction modulo* (Dowek et al., 1998), that allows to identify propositions modulo a congruence.

In deduction modulo, the notions of language, term and proposition are that of many-sorted first-order logic. We consider *theories* formed with a set of axioms Γ *and a congruence*, denoted $\equiv$, defined on propositions. The deduction rules take into account this

equivalence, for instance, the right rule of the conjunction is not given as usual:

$$\frac{\Gamma \vdash A, \Delta \quad \Gamma \vdash B, \Delta}{\Gamma \vdash A \wedge B, \Delta}$$

but it is formulated as:

$$\frac{\Gamma \vdash A, \Delta \quad \Gamma \vdash B, \Delta}{\Gamma \vdash C, \Delta} \text{ if } C \equiv A \wedge B$$

and all the rules of sequent calculus are stated in a similar way as described in figure 1.

When Γ is a finite set of axioms and $\equiv$ a congruence, a proposition P is said to be provable in $(\Gamma, \equiv)$ if the sequent $\Gamma \vdash P$ is derivable modulo $\equiv$. When Γ is infinite, a proposition P is said to be provable in $(\Gamma, \equiv)$ if it is provable in $(\Gamma', \equiv)$, where Γ' is a finite subset of Γ.

For instance, in sequent calculus modulo the congruence defined by the rewrite system:

$$\begin{array}{rcl} 0 + y & \longrightarrow & y \\ S(x) + y & \longrightarrow & S(x + y) \\ 0 \times y & \longrightarrow & 0 \\ S(x) \times y & \longrightarrow & x \times y + y \end{array}$$

we can prove that the number 4 is even.

$$\dfrac{\dfrac{\dfrac{}{4 = 4 \vdash 2 \times 2 = 4}\,\text{axiom}}{\forall x\ x = x \vdash 2 \times 2 = 4}\,(x, x = x, 4)\ \forall\text{-l}}{\forall x\ x = x \vdash \exists y\ 2 \times y = 4}\,(y, 2 \times y = 4, 2)\ \exists\text{-r}$$

Substituting the variable y by the term 2 in the proposition $2 \times y = 4$, as indicated on the side of the rule, yields the proposition $2 \times 2 = 4$, that is congruent to $4 = 4$. The transformation of one proposition into the other, that would require several proof steps in sequent calculus using the axioms of arithmetic, does not appear here. Since the congruence is decidable, this computation needs not to be recorded in the proof itself.

Notice that we do not use the axioms of addition and multiplication explicitly in the proof. Indeed, these axioms are now redundant: as the terms $0 + y$ and y are congruent, the axiom $\forall y\ 0 + y = y$ is congruent to the equality axiom $\forall y\ y = y$. Hence, it can be dropped. Using the terminology introduced by Plotkin, these axioms have been *built-in* (Plotkin, 1972; Andrews, 1971; Peterson and Stickel, 1981; Stickel, 1985; Jouannaud and Kirchner, 1986; Marché, 1994; Viry, 1995; Viry, 1998).

In the example above, the congruence is just the congruent closure of the relation induced on terms by the term rewriting system. In many situations, it is also natural to consider congruences defined directly at the proposition level. For instance, we may add to the previous system the rule of integral domains

$$x \times y = 0 \longrightarrow x = 0 \vee y = 0$$

that rewrites an atomic proposition to a disjunction. As far as we know, this property cannot be expressed by term rewriting rules. Deduction modulo can consider congruences

$$\frac{}{P \vdash Q}\text{axiom} \;\; \text{if } P \equiv Q \qquad \frac{\Gamma, P \vdash \Delta \quad \Gamma \vdash Q, \Delta}{\Gamma \vdash \Delta}\text{cut} \;\; \text{if } P \equiv Q$$

$$\frac{\Gamma, Q_1, Q_2 \vdash \Delta}{\Gamma, P \vdash \Delta}\text{contr-l} \;\; \text{if } P \equiv Q_1 \equiv Q_2 \qquad \frac{\Gamma \vdash Q_1, Q_2, \Delta}{\Gamma \vdash P, \Delta}\text{contr-r} \;\; \text{if } P \equiv Q_1 \equiv Q_2$$

$$\frac{\Gamma \vdash \Delta}{\Gamma, P \vdash \Delta}\text{weak-l} \qquad \frac{\Gamma \vdash \Delta}{\Gamma \vdash P, \Delta}\text{weak-r}$$

$$\frac{\Gamma \vdash P, \Delta \quad \Gamma, Q \vdash \Delta}{\Gamma, R \vdash \Delta}\Rightarrow\text{-l} \;\; \text{if } R \equiv (P \Rightarrow Q) \qquad \frac{P, \Gamma \vdash Q, \Delta}{\Gamma \vdash R, \Delta}\Rightarrow\text{-r} \;\; \text{if } R \equiv (P \Rightarrow Q)$$

$$\frac{\Gamma, P, Q \vdash \Delta}{\Gamma, R \vdash \Delta}\wedge\text{-l} \;\; \text{if } R \equiv (P \wedge Q) \qquad \frac{\Gamma \vdash P, \Delta \quad \Gamma \vdash Q, \Delta}{\Gamma \vdash R, \Delta}\wedge\text{-r} \;\; \text{if } R \equiv (P \wedge Q)$$

$$\frac{\Gamma, P \vdash \Delta \quad \Gamma, Q \vdash \Delta}{\Gamma, R \vdash \Delta}\vee\text{-l} \;\; \text{if } R \equiv (P \vee Q) \qquad \frac{\Gamma \vdash P, Q, \Delta}{\Gamma \vdash R, \Delta}\vee\text{-r} \;\; \text{if } R \equiv (P \vee Q)$$

$$\frac{\Gamma \vdash P, \Delta}{\Gamma, R \vdash \Delta}\neg\text{-l} \;\; \text{if } R \equiv \neg P \qquad \frac{\Gamma, P \vdash \Delta}{\Gamma \vdash R, \Delta}\neg\text{-r} \;\; \text{if } R \equiv \neg P$$

$$\frac{}{\Gamma, P \vdash \Delta}\bot\text{-l} \;\; \text{if } P \equiv \bot$$

$$\frac{\Gamma, \{t/x\}P \vdash \Delta}{\Gamma, Q \vdash \Delta}(x, P, t)\ \forall\text{-l} \;\; \text{if } Q \equiv \forall x\ P \qquad \frac{\Gamma \vdash \{y/x\}P, \Delta}{\Gamma \vdash Q, \Delta}(x, P, y)\ \forall\text{-r} \;\; \text{if } Q \equiv \forall x\ P$$

$$\frac{\Gamma, \{y/x\}P \vdash \Delta}{\Gamma, Q \vdash \Delta}(x, P, y)\ \exists\text{-l} \;\; \text{if } Q \equiv \exists x\ P \qquad \frac{\Gamma \vdash \{t/x\}P, \Delta}{\Gamma \vdash Q, \Delta}(x, P, t)\ \exists\text{-r} \;\; \text{if } Q \equiv \exists x\ P$$

where the rules $\forall$-r and $\exists$-l assume that $y \notin FV(\Gamma\Delta)$

Fig. 1. The sequent calculus modulo

on propositions defined by rules rewriting terms to terms and atomic propositions to arbitrary ones. Rules with non-atomic left-hand sides — that are technically much more difficult to handle — are not tackled in this work.

All congruences in this paper are defined by confluent rewrite systems. As these rewrite systems are defined on propositions and propositions contain binders, these rewrite systems are in fact *Combinatory Reduction Systems* (Klop et al., 1993).

Notice that deduction modulo is not a proper extension of first-order logic. It is proved in (Dowek et al., 1998) that for every congruence $\equiv$, we can find a theory $\mathcal{T}$ such that $\Gamma \vdash P$ is provable modulo $\equiv$ if and only if $\mathcal{T}, \Gamma \vdash P$ is provable in ordinary first-order logic. Of course, the provable propositions are the same, but the proofs are very different, indeed much shorter in deduction modulo.

Proof search in deduction modulo can be handled by a method called *Extended Narrowing and Resolution* (ENAR) that extends the usual resolution method and that is described below.

2. HOL-λ

We recall quickly the usual presentation of higher-order logic. Terms are those of a simply typed λ-calculus (Girard et al., 1989) with two base types ι and o and the following constants:

— $\dot{\Rightarrow}$, $\dot{\wedge}$ and $\dot{\forall}$ of type $o \to o \to o$,
— $\dot{\neg}$ of type $o \to o$,
— $\dot{\bot}$ of type o,
— for each type T constants, $\dot{\forall}_T$ and $\dot{\exists}_T$ of type $(T \to o) \to o$.

The notation with a dot for the constants let us distinguish them from the connectors and quantifiers of first-order logic. Terms of type o are called *propositions*.

The unique $\beta\eta$-normal form of a term a is written $a\downarrow$. The deduction rules of HOL-λ are given in figure 2 where all propositions are supposed to be $\beta\eta$-normal.

For instance, we can prove the sequent

$$(\dot{\forall}\ \lambda P\ (\dot{\Rightarrow}\ (P\ a)\ (P\ b))), (R\ a\ a) \vdash (R\ b\ b)$$

as follows.

$$\dfrac{\dfrac{\dfrac{}{(R\ a\ a) \vdash (R\ a\ a)}\ \text{axiom} \qquad \dfrac{}{(R\ b\ b) \vdash (R\ b\ b)}\ \text{axiom}}{(\dot{\Rightarrow}\ (R\ a\ a)\ (R\ b\ b)), (R\ a\ a) \vdash (R\ b\ b)}\ \dot{\Rightarrow}\text{-l}}{(\dot{\forall}\ \lambda P\ (\dot{\Rightarrow}\ (P\ a)\ (P\ b))), (R\ a\ a) \vdash (R\ b\ b)}\ \dot{\forall}\text{-l}$$

where in the $\dot{\forall}$-left rule, the term $\lambda P\ (\dot{\Rightarrow}\ (P\ a)\ (P\ b))$ has been applied to the term $\lambda x\ (R\ x\ x)$.

In an alternative presentation of HOL-λ, propositions are not normalized in the quantifier rules. Instead, such a presentation takes axioms stating that two $\beta\eta$-convertible terms are equal, see for instance (Church, 1940; Andrews, 1986). In this case, for the example above, applying the term $\lambda P\ (\dot{\Rightarrow}\ (P\ a)\ (P\ b))$ to the term $\lambda x\ (R\ x\ x)$ we get $(\lambda P\ (\dot{\Rightarrow}\ (P\ a)\ (P\ b))\ \lambda x\ (R\ x\ x))$ and we use the axioms to deduce $(\dot{\Rightarrow}\ (R\ a\ a)\ (R\ b\ b))$.

The system HOL-λ is well-known to be consistent and to enjoy cut elimination (Girard, 1970; Girard, 1972).

In HOL-λ, equality needs not to be primitive. It can be defined as Leibniz' equality i.e. $\lambda x\ \lambda y\ \dot{\forall}\lambda p\ ((p\ x)\dot{\Rightarrow}(p\ y))$. In this case, the propositions

$$\dot{\forall}\ \lambda f\ \dot{\forall}\ \lambda g\ \dot{\forall}\ \lambda x\ (f = g \dot{\Rightarrow} (f\ x) = (g\ x))$$
$$\dot{\forall}\ \lambda x\ \dot{\forall}\ \lambda y\ \dot{\forall}\ \lambda f\ (x = y \dot{\Rightarrow} (f\ x) = (f\ y))$$

are provable. But the proposition

$$\dot{\forall}\ \lambda f\ \dot{\forall}\ \lambda g\ (\dot{\forall}\ \lambda x\ ((f\ x) = (g\ x)) \dot{\Rightarrow} \lambda x\ (f\ x) = \lambda x\ (g\ x))$$

is not because substitutions avoid captures. Similarly, the propositions

$$\dot{\forall}\ \lambda f\ \dot{\forall}\ \lambda g\ ((\dot{\forall}\ \lambda x\ ((f\ x) = (g\ x))) \dot{\Rightarrow} f = g)$$

and

$$\dot{\forall}\ \lambda x\ \dot{\forall}\ \lambda y\ ((x \dot{\Leftrightarrow} y) \dot{\Rightarrow} x = y)$$

$$\frac{}{P \vdash P}\text{axiom} \qquad \frac{\Gamma, P \vdash \Delta \quad \Gamma \vdash P, \Delta}{\Gamma \vdash \Delta}\text{cut}$$

$$\frac{\Gamma, P, P \vdash \Delta}{\Gamma, P \vdash \Delta}\text{contr-l} \qquad \frac{\Gamma \vdash P, P, \Delta}{\Gamma \vdash P, \Delta}\text{contr-r}$$

$$\frac{\Gamma \vdash \Delta}{\Gamma, P \vdash \Delta}\text{weak-l} \qquad \frac{\Gamma \vdash \Delta}{\Gamma \vdash P, \Delta}\text{weak-r}$$

$$\frac{\Gamma \vdash P, \Delta \quad \Gamma, Q \vdash \Delta}{\Gamma, (\dot{\Rightarrow}\ P\ Q) \vdash \Delta}\dot{\Rightarrow}\text{-l} \qquad \frac{P, \Gamma \vdash Q, \Delta}{\Gamma \vdash (\dot{\Rightarrow}\ P\ Q), \Delta}\dot{\Rightarrow}\text{-r}$$

$$\frac{\Gamma, P, Q \vdash \Delta}{\Gamma, (\dot{\wedge}\ P\ Q) \vdash \Delta}\dot{\wedge}\text{-l} \qquad \frac{\Gamma \vdash P, \Delta \quad \Gamma \vdash Q, \Delta}{\Gamma \vdash (\dot{\wedge}\ P\ Q), \Delta}\dot{\wedge}\text{-r}$$

$$\frac{\Gamma, P \vdash \Delta \quad \Gamma, Q \vdash \Delta}{\Gamma, (\dot{\vee}\ P\ Q) \vdash \Delta}\dot{\vee}\text{-l} \qquad \frac{\Gamma \vdash P, Q, \Delta}{\Gamma \vdash (\dot{\vee}\ P\ Q), \Delta}\dot{\vee}\text{-r}$$

$$\frac{\Gamma \vdash P, \Delta}{\Gamma, (\dot{\neg}\ P) \vdash \Delta}\dot{\neg}\text{-l} \qquad \frac{\Gamma, P \vdash \Delta}{\Gamma \vdash (\dot{\neg}\ P), \Delta}\dot{\neg}\text{-r}$$

$$\frac{}{\Gamma, \dot{\bot} \vdash \Delta}\dot{\bot}\text{-l}$$

$$\frac{\Gamma, (P\ t)\downarrow \vdash \Delta}{\Gamma, (\dot{\forall}\ P) \vdash \Delta}t\ \dot{\forall}\text{-l} \qquad \frac{\Gamma \vdash (P\ y)\downarrow, \Delta}{\Gamma \vdash (\dot{\forall}\ P), \Delta}\dot{\forall}\text{-r}$$

$$\frac{\Gamma, (P\ y)\downarrow \vdash \Delta}{\Gamma, (\dot{\exists}\ P) \vdash \Delta}\dot{\exists}\text{-l} \qquad \frac{\Gamma \vdash (P\ t)\downarrow, \Delta}{\Gamma \vdash (\dot{\exists}\ P), \Delta}t\ \dot{\exists}\text{-r}$$

where the rules $\dot{\forall}$-r and $\dot{\exists}$-l assume that $y \notin FV(\Gamma\Delta)$

Fig. 2. **HOL-** : The deduction rules of HOL-λ

are not provable.

These two last propositions can be added as axioms in the theory. They are called *extensionality axioms* and the theory itself is called *extensional higher-order logic*. In contrast, without these extensionality axioms, the theory is called *intentional higher-order logic*.

3. HOL-C

3.1. *HOL-C as a first-order theory*

Higher-order logic can be expressed as a many-sorted first-order theory with equality. The sorts of this theory are the types of simply typed λ-calculus, i.e. they are inductively defined by:

— ι and o are sorts,
— if T and U are sorts then $T \to U$ is a sort.

A function symbol f is said to have *rank* $(T_1, \ldots, T_n)\ U$ if it takes as arguments n terms of sorts $T_1, \ldots, T_n$ and constructs a term of sort U. A predicate symbol P is said to have *rank* $(T_1, \ldots, T_n)$ if it takes as arguments n terms of sorts $T_1, \ldots, T_n$.

Besides equality, the language contains the function symbols:

— $\alpha_{T,U}$ of rank $(T \to U, T)\ U$

These symbols are called *applications*. As usual, the term $\alpha_{T,U}(t, u)$ is written $(t\ u)$ and $(\ldots(t\ u_1)\ldots u_n)$ is written $(t\ u_1\ \ldots\ u_n)$.

Then, the language contains the unary predicate symbol:

— ε of rank (o)

that transforms a term t of sort o into the proposition $\varepsilon(t)$.

To express function terms and predicate terms, instead of using λ-calculus, we introduce for each n-tuple of variables $x_1, \ldots, x_n$ respectively of sorts $T_1, \ldots, T_n$ and each term t of sort U formed with the variables $x_1, \ldots, x_n$ and the application symbols, a constant symbol:

— $x_1, \ldots, x_n \longmapsto t$ of sort $T_1 \to \ldots \to T_n \to U$

Such constant symbols are called *combinators*. Finally, the language contains also the constant symbols:

— $\dot{\Rightarrow}$, $\dot{\wedge}$ and $\dot{\vee}$ of sort $o \to o \to o$,
— $\dot{\neg}$ of sort $o \to o$,
— $\dot{\bot}$ of sort o,
— $\dot{\forall}_T$ and $\dot{\exists}_T$ of sort $(T \to o) \to o$.

Besides the well-known axioms of equality, the theory contains the following axioms that express the meaning of the combinators:

$$((x_1, \ldots, x_n \longmapsto t)\ x_1\ \ldots\ x_n) = t$$

and axioms that relate the connectors and quantifiers (e.g. $\wedge$) and their replication as constant symbols (e.g. $\dot{\wedge}$):

$$\begin{aligned}
\varepsilon(\dot{\Rightarrow}\ x\ y) &\Leftrightarrow (\varepsilon(x) \Rightarrow \varepsilon(y)) \\
\varepsilon(\dot{\wedge}\ x\ y) &\Leftrightarrow (\varepsilon(x) \wedge \varepsilon(y)) \\
\varepsilon(\dot{\vee}\ x\ y) &\Leftrightarrow (\varepsilon(x) \vee \varepsilon(y)) \\
\varepsilon(\dot{\neg}\ x) &\Leftrightarrow \neg\varepsilon(x) \\
\varepsilon(\dot{\bot}) &\Leftrightarrow \bot \\
\varepsilon(\dot{\forall}\ x) &\Leftrightarrow \forall y\ \varepsilon(x\ y) \\
\varepsilon(\dot{\exists}\ x) &\Leftrightarrow \exists y\ \varepsilon(x\ y)
\end{aligned}$$

Notice that, with our convention on applications notation, the term $(\dot{\Rightarrow}\ x\ y)$ is indeed $\alpha(\alpha(\dot{\Rightarrow}, x), y)$ and that $\varepsilon(\dot{\Rightarrow}\ x\ y)$ is an atomic proposition.

3.2. *HOL-C as a first-order theory modulo*

In deduction modulo, these axioms can be built-in. So, we work modulo the congruence defined by the rewriting system $\mathcal{R}$ containing the following term rewrite rules:

$$((x_1, \ldots, x_n \longmapsto t)\ u_1\ \ldots\ u_n) \longrightarrow t\{x_1 \mapsto u_1, \ldots, x_n \mapsto u_n\}$$

and the following proposition rewrite rules:

$$\begin{aligned}
\varepsilon(\dot{\Rightarrow}\ x\ y) &\longrightarrow \varepsilon(x) \Rightarrow \varepsilon(y)\\
\varepsilon(\dot{\wedge}\ x\ y) &\longrightarrow \varepsilon(x) \wedge \varepsilon(y)\\
\varepsilon(\dot{\vee}\ x\ y) &\longrightarrow \varepsilon(x) \vee \varepsilon(y)\\
\varepsilon(\dot{\neg}\ x) &\longrightarrow \neg\varepsilon(x)\\
\varepsilon(\dot{\bot}) &\longrightarrow \bot\\
\varepsilon(\dot{\forall}\ x) &\longrightarrow \forall y\ \varepsilon(x\ y)\\
\varepsilon(\dot{\exists}\ x) &\longrightarrow \exists y\ \varepsilon(x\ y)
\end{aligned}$$

Notice that the rules of the first group rewrite terms, while the rules of the second group rewrite atomic propositions. Hence, we have here a typical example where rules rewriting terms are not enough.

Remark also that in this formulation, equality can be defined as Leibniz' equality. Indeed, if a and b are terms of type T, the notation $a = b$ may be introduced as an abbreviation for the proposition $\forall p\ (\varepsilon(p\ a) \Rightarrow \varepsilon(p\ b))$, where p is a variable of type $T \to o$. Thus there is no need to take equality as a primitive symbol.

3.3. *The extensionality axioms*

In HOL-C, the extensionality axioms are

$$\forall f\ \forall g\ ((\forall x\ ((f\ x) = (g\ x))) \Rightarrow f = g)$$

$$\forall x\ \forall y\ (\varepsilon(x) \Leftrightarrow \varepsilon(y)) \Rightarrow x = y)$$

Here equality is either a primitive symbol as in section 3.1 or an abbreviation as in section 3.2.

3.4. *Embedding HOL-λ into HOL-C*

The translation from λ-terms to combinators is usually called λ-*lifting* and is denoted by $(\_)_C$. Applications, variables and constants are translated in an obvious way to their correspondents in HOL-C. A term of the form $\lambda x\ t$ is translated as follows. Let $y_1, \ldots, y_n$ be the variables of t_C minus x. Then, in t_C, all the occurrences of any combinator c_i are replaced by a fresh variable z_i yielding a term $(t_C)'$. Then,

$$(\lambda x\ t)_C = ((y_1, \ldots, y_n, z_1, \ldots, z_p, x \longmapsto (t_C)')\ y_1\ \ldots\ y_n\ c_1\ \ldots\ c_p)$$

For instance the term

$$((\lambda x\ \lambda y\ x)\ (u\ v))$$

is translated as

$$((f, x \longmapsto (f\ x))\ (x, y \longmapsto x)\ (u\ v)).$$

This translation can be modified in order to use only the combinators $S = x, y, z \longmapsto ((x\ z)\ (y\ z))$ and $K = x, y \longmapsto x$.

Extensional HOL-C can be shown to be equivalent to extensional HOL-λ, i.e. a proposition P is provable in extensional HOL-λ if and only if the proposition $\varepsilon(P_C)$ is provable in extensional HOL-C.

But, if we drop the extensionality axioms, then the two presentations are not equivalent anymore. For instance, the proposition

$$((\lambda x\ \lambda y\ x)\ (u\ v)) = \lambda y\ (u\ v)$$

is provable in HOL-λ while its translation in HOL-C

$$\varepsilon(((f, x \longmapsto (f\ x))\ (x, y \longmapsto x)\ (u\ v)) = ((u, v, y \longmapsto (u\ v))\ u\ v))$$

requires extensionality.

Even for extensional higher-order logic, the formulations with λ-calculus and combinators are only weakly equivalent: provable propositions are the same, but the proofs are very different, since some proofs requiring only $\beta\eta$-conversion in HOL-λ require the use of extensionality in HOL-C.

4. HOL-$\lambda\sigma$

We present now the definition and first properties of the first-order theory modulo HOL-$\lambda\sigma$, and we will see in the next section that it provides a new first-order formulation of higher-order logic. The theory HOL-$\lambda\sigma$ is not based on combinators as previously, but on de Bruijn indices and explicit substitutions. It allows to avoid the drawbacks of the formulation with combinators mentioned just above.

4.1. *HOL-λσ as a first-order theory modulo*

In λ-calculus with de Bruijn indices, bound variables are replaced by an index indicating the binding height of this variable, i.e. the number of λ's between this occurrence and its binder. For instance the term $\lambda x\ (x\ (\lambda y\ x))$ is written $\lambda\ (1\ (\lambda\ 2))$. So λ-calculus with de Bruijn notation is also a first-order language with a binary function symbol α, a unary function symbol λ and constant symbols $1, 2, 3 \ldots$.

Types of simply typed λ-calculus are not sufficient anymore with de Bruijn indices. Indeed, we need to give a sort not only to terms like $(\lambda_T\ 1)$ (that gets the sort $T \to T$), but also to terms of the form 1. Thus, as detailed in (Dowek et al., 1995), we have to consider sorts of the form $\Gamma \vdash T$ where T is a type of simply typed λ-calculus and Γ a context, i.e. a list of such types. For example, the term 1 has the sort $A.\Gamma \vdash A$.

With de Bruijn indices only, conversion axioms use an external definition for substitution. Moreover this substitution is not well-defined on open terms of this first-order language. This is solved by considering an extension of this calculus: the *calculus of explicit substitutions* (Abadi et al., 1991) also called $\lambda\sigma$-calculus. Beside the sorts of the form $\Gamma \vdash T$, this calculus introduces also sorts of the form $\Gamma \vdash \Delta$ for substitutions that are lists of terms. The symbols to build such substitutions are id, ., $\uparrow$ and $\circ$. Then, a new term constructor is introduced $\_[\_]$ that allows to apply an explicit substitution to a term. The rewrite rules describing the evaluation of the $\lambda\sigma$-calculus are given in figure

β-reduction and η-reduction:

$$\begin{aligned}(\lambda a)b &\longrightarrow a[b.id]\\ \lambda(a\ 1) &\longrightarrow b \text{ if } a =_\sigma b[\uparrow]\end{aligned}$$

σ-reduction:

$$\begin{aligned}(a\ b)[s] &\longrightarrow (a[s]\ b[s])\\ \mathtt{1}[a.s] &\longrightarrow a\\ a[id] &\longrightarrow a\\ (\lambda a)[s] &\longrightarrow \lambda(a[\mathtt{1}.(s\circ\uparrow)])\\ (a[s])[t] &\longrightarrow a[s\circ t]\\ id\circ s &\longrightarrow s\\ \uparrow\circ(a.s) &\longrightarrow s\\ (s_1\circ s_2)\circ s_3 &\longrightarrow s_1\circ(s_2\circ s_3)\\ (a.s)\circ t &\longrightarrow a[t].(s\circ t)\\ s\circ id &\longrightarrow s\\ \mathtt{1}.\uparrow &\longrightarrow id\\ \mathtt{1}[s].(\uparrow\circ s) &\longrightarrow s\end{aligned}$$

Fig. 3. The rewrite rules of $\lambda\sigma$-calculus

3. They will be used to define (a part of) the congruence, so they give an example where the congruence is defined from a conditional rewrite system.

Now we introduce HOL-$\lambda\sigma$. It is a many-sorted first-order theory modulo with sorts of the form $\Gamma \vdash T$ and $\Gamma \vdash \Delta$ where Γ and Δ are sequences of types of simply typed λ-calculus and T is such a type.

Definition 4.1. (Language of HOL-$\lambda\sigma$) The language contains the following function symbols:

$\mathtt{1}^\Gamma_A$	constant of sort	$A.\Gamma \vdash A$
$\alpha^\Gamma_{A\to B,A}$	binary function of rank	$(\Gamma \vdash A \to B, \Gamma \vdash A)\Gamma \vdash B$
$\lambda^\Gamma_{A,B}$	unary function of rank	$(A.\Gamma \vdash B)\Gamma \vdash A \to B$
$[\ ]^{\Gamma,\Gamma'}_A$	binary function of rank	$(\Gamma' \vdash A, \Gamma \vdash \Gamma')\Gamma \vdash A$
id^Γ	constant of sort	$\Gamma \vdash \Gamma$
$\uparrow^\Gamma_A$	constant of sort	$A.\Gamma \vdash \Gamma$
$\cdot^{\Gamma,\Gamma'}_A$	binary function of rank	$(\Gamma \vdash A, \Gamma \vdash \Gamma')\Gamma \vdash A.\Gamma'$
$\circ^{\Gamma,\Gamma',\Gamma''}$	binary function of rank	$(\Gamma \vdash \Gamma'', \Gamma'' \vdash \Gamma')\Gamma \vdash \Gamma'$
$\dot\Rightarrow$	constant of sort	$\vdash o \to o \to o$
$\dot\wedge$	constant of sort	$\vdash o \to o \to o$
$\dot\vee$	constant of sort	$\vdash o \to o \to o$
$\dot\neg$	constant of sort	$\vdash o \to o$
$\dot\bot$	constant of sort	$\vdash o$
$\dot\forall_A$	constant of sort	$\vdash (A \to o) \to o$
$\dot\exists_A$	constant of sort	$\vdash (A \to o) \to o$

and a single unary predicate symbol:

$$\begin{array}{rcl}
\varepsilon(\dot{\Rightarrow}\ x\ y) & \longrightarrow & \varepsilon(x) \Rightarrow \varepsilon(y) \\
\varepsilon(\dot{\wedge}\ x\ y) & \longrightarrow & \varepsilon(x) \wedge \varepsilon(y) \\
\varepsilon(\dot{\vee}\ x\ y) & \longrightarrow & \varepsilon(x) \vee \varepsilon(y) \\
\varepsilon(\dot{\neg}\ x) & \longrightarrow & \neg\varepsilon(x) \\
\varepsilon(\dot{\bot}) & \longrightarrow & \bot \\
\varepsilon(\dot{\forall}_T\ x) & \longrightarrow & \forall y\ \varepsilon(x\ y) \\
\varepsilon(\dot{\exists}_T\ x) & \longrightarrow & \exists y\ \varepsilon(x\ y)
\end{array}$$

Fig. 4. The $\mathcal{L}$-rewrite rules

$$\varepsilon \quad \text{of rank} \quad (\vdash o)$$

We denote $\lambda\sigma\mathcal{L}$ the rewrite rules of $\lambda\sigma$-calculus together with the logical rules $\mathcal{L}$ given in figure 4 and we write $A \equiv_{\lambda\sigma\mathcal{L}} B$ when A and B are congruent modulo $\lambda\sigma\mathcal{L}$.

4.2. *HOL-λσ as a first-order theory*

We have presented HOL-$\lambda\sigma$ as a first-order theory modulo and it will be used latter that way to represent HOL-λ. We have proved in (Dowek et al., 1998) that any theory modulo $(\Gamma, \equiv)$ could be also expressed as a (non modulo) theory, i.e. that there exists a set of axioms $\mathcal{T}$ such that $\Gamma \vdash P$ modulo $\equiv$ if and only if $\mathcal{T}, \Gamma \vdash P$ in standard first-order logic.

As any theory modulo, HOL-$\lambda\sigma$ can be expressed as a first-order theory. The naive expression takes as axioms the universal closures of all propositions of the form $P \Leftrightarrow Q$ where $P \equiv_{\lambda\sigma\mathcal{L}} Q$.

In (Dowek et al., 1998), we have also shown that when the congruence is defined by rewrite rules, we can take less axioms: we first take an equality predicate and the axioms of equality, then for each rewrite rule $l \to r$, we take as axiom the universal closure of the proposition $l = r$ when l and r are terms or $l \Leftrightarrow r$ when l and r are propositions. This result does not apply here because the rule η is a conditionnal rewrite rule. However, for this rule we can take the axiom

$$\forall x\ (\lambda(x[\uparrow]\ 1) = x)$$

and we get a presentation of HOL-$\lambda\sigma$ as a (non modulo) first-order theory.

4.3. *Properties of HOL-λσ*

Proposition 4.1. (Termination) The system $\lambda\sigma\mathcal{L}$ is weakly terminating.

Proof. Since the $\mathcal{L}$ and $\lambda\sigma$ rewrite systems share the application operator α, we cannot try to apply the existing termination modularity results.

Typed $\lambda\sigma$-calculus may not terminate (Melliès, 1995), but it is known that the strategy σ-normalizing the term after each application of β or η (Goubault-Larrecq, 1997; Muñoz, 1997a) is normalizing. We reduce termination for $\lambda\sigma\mathcal{L}$ to termination for this normalizing strategy of $\lambda\sigma$. So, we define a translation of the terms and the propositions of HOL-$\lambda\sigma$

into the typed system $\lambda\sigma$, denoted by $||\_||$, as follows. In each sort $\Gamma \vdash T$, we choose a variable $z_{\Gamma\vdash T}$.

— $||x|| = z_T$, where x is a variable,
— every symbol is mapped to itself but:
  – $||\dot{\Rightarrow}|| = ||\dot{\wedge}|| = ||\dot{\vee}|| = ((\lambda 1)\ z_{\vdash o\to o\to o})$,
  – $||\dot{\neg}|| = (\lambda 1)$,
  – $||\dot{\bot}|| = ((\lambda 1)\ z_{\vdash o})$,
  – $||\dot{\forall}_T|| = ||\dot{\exists}_T|| = \lambda(1\ z_{\vdash T}[\uparrow])$,
— $||\varepsilon(t)|| = ||t||$,
— $||P \Rightarrow Q|| = ||P \wedge Q|| = ||P \vee Q|| = (z_{\vdash o\to o\to o}\ ||P||\ ||Q||)$,
— $||\neg P|| = ||P||$,
— $||\bot|| = z_o$,
— $||\forall x\ P|| = ||\exists x\ P|| = ||P||$.

In $\lambda\sigma\mathcal{L}$, we say that t R_1-reduces to u if u is obtained by reducing a β-redex, a η-redex or a $\mathcal{L}$-redex and σ-normalizing the term obtained this way. In $\lambda\sigma$, we say that t R_2-reduces to u if u is obtained by reducing a β-redex or a η-redex and σ-normalizing the term obtained this way. We check that if P R_1-rewrites in one step to Q, then $||P||$ R_2-rewrites in at least one step to $||Q||$. Let $P_1, P_2, \ldots$ be a R_1-reduction sequence in the above system, the sequence $||P_1||, ||P_2||, \ldots$ is a R_2-reduction sequence in $\lambda\sigma$, thus it is finite. □

Proposition 4.2. (Confluence) $\lambda\sigma\mathcal{L}$ is confluent on terms containing only term variables.

Proof. Since the $\mathcal{L}$ and $\lambda\sigma$ rewrite systems share the application operator α, we cannot apply Toyama's modularity result.

The proof is based on Hindley-Rosen lemma (Hindley, 1964; Rosen, 1973): if two relations R and S are strongly confluent (i.e. if $t\ R\ u$ and $t\ R\ v$ then there exists a w such that $u\ R\ w$ and $v\ R\ w$ and if $t\ S\ u$ and $t\ S\ v$ then there exists a w such that $u\ S\ w$ and $v\ S\ w$) and strongly commute (i.e. if $t\ R\ u$ and $t\ S\ v$ then there exists a term w such that $u\ S\ w$ and $v\ R\ w$) then the relation $R \cup S$ is confluent.

For R we take $\mathcal{L}$ and for S we take $\lambda\sigma^*$. The system $\mathcal{L}$ is linear and orthogonal, hence it is strongly confluent. Since $\lambda\sigma$ is confluent (Abadi et al., 1991) the rewrite relation $\lambda\sigma^*$ is strongly confluent. Finally, $\mathcal{L}$ and $\lambda\sigma^*$ strongly commute. Indeed, if $(t\ \mathcal{L}\ u)$ and $(t\ \lambda\sigma^*\ v)$ then the $\mathcal{L}$-redex in t is either disjoint from or above the $\lambda\sigma$ redex. In both cases we can reduce the $\lambda\sigma$-redex in u and the $\mathcal{L}$ redex in v getting the same term. Hence, $\lambda\sigma\mathcal{L}$ is confluent. □

Proposition 4.3. (Consistency) The theory HOL-$\lambda\sigma$ is consistent.

Proof. We construct a model as follows:

— $\mathcal{M}_\iota = \{0\}$,
— $\mathcal{M}_o = \{0, 1\}$,
— $\mathcal{M}_{T\to U} = \mathcal{M}_U^{\mathcal{M}_T}$,
— $\mathcal{D}_{T_1,\ldots,T_n\vdash U} = \mathcal{M}_U^{\mathcal{M}_{T_n}^{\cdots^{\mathcal{M}_{T_1}}}}$,

— $\mathcal{D}_{T_1,\ldots,T_n \vdash U_1,\ldots,U_p} = (\mathcal{M}_{U_1} \times \ldots \times \mathcal{M}_{U_p})^{\mathcal{M}_{T_n}\ldots^{\mathcal{M}_{T_1}}}$.

If f is a function of the set $(\mathcal{M}_{U_1} \times \ldots \times \mathcal{M}_{U_p})^{\mathcal{M}_{T_n}\ldots^{\mathcal{M}_{T_1}}}$ and g a function of the set $(\mathcal{M}_{V_1} \times \ldots \times \mathcal{M}_{V_q})^{\mathcal{M}_{U_p}\ldots^{\mathcal{M}_{U_1}}}$ we write $g \circ f$ for the function of $(\mathcal{M}_{V_1} \times \ldots \times \mathcal{M}_{V_q})^{\mathcal{M}_{T_n}\ldots^{\mathcal{M}_{T_1}}}$ mapping $x_1, \ldots, x_n$ to $g(t_1)\ldots(t_p)$ where $(t_1, \ldots, t_p) = f(x_1)\ldots(x_n)$

Then, we interpret the symbols of the language as follows.

— $\overline{1_A^\Gamma}$ is the function mapping $a_1, \ldots, a_n$ to a_1.

— $\overline{\alpha^\Gamma_{A\to B,A}}$ is the function mapping a and b to the function mapping $c_1, \ldots, c_n$ to $a(c_1, \ldots, c_n)(b(c_1, \ldots, c_n))$.

— $\overline{\lambda^\Gamma_{A,B}}$ is the identity function.

— $\overline{[\,]_A^{\Gamma,\Gamma'}}$ is the function mapping a and b to $b \circ a$.

— $\overline{id^\Gamma}$ is the identity function.

— $\overline{\uparrow_A^\Gamma}$ is the function mapping $a_1, \ldots, a_n$ to $(a_2, \ldots, a_n)$.

— $\overline{\cdot_A^{\Gamma,\Gamma'}}$ is the function mapping a, b to the function mapping $c_1, \ldots, c_n$ to $(a(c_1, \ldots, c_n), b(c_1, \ldots, c_n))$.

— $\overline{\circ^{\Gamma,\Gamma',\Gamma''}}$ is the function mapping a and b to $b \circ a$.

— $\overline{\dot{\Rightarrow}}$ is the function mapping a and b to 1 if $a = 0$ or $b = 1$ and to 0 otherwise.

— $\overline{\dot{\wedge}}$ maps a and b to 1 if $a = 1$ and $b = 1$ and to 0 otherwise.

— $\overline{\dot{\vee}}$ maps a and b to 1 if $a = 1$ or $b = 1$ and to 0 otherwise.

— $\overline{\dot{\neg}}$ maps a to 1 if $a = 0$ and to 0 otherwise.

— $\overline{\dot{\bot}} = 0$.

— $\overline{\dot{\forall}_T}$ maps a to 1 if a maps every object of $\mathcal{M}_T$ to 1 and to 0 otherwise.

— $\overline{\dot{\exists}_T}$ maps a to 1 if a maps some object of $\mathcal{M}_T$ to 1 and to 0 otherwise.

— $\overline{\varepsilon}$ is the identity function.

We check that if $A \equiv_{\lambda\sigma\mathcal{L}} B$ then A and B have the same denotation. Then we check that every provable proposition denotes the truth value 1 and hence that $\bot$ is not derivable. □

We show now that any proof in HOL-$\lambda\sigma$ can be transformed into a cut free proof, i.e. a proof built without the rule *cut*. This result will be used twice, to show the equivalence between HOL-λ and HOL-$\lambda\sigma$ and to prove the completeness of the ENAR method for HOL-$\lambda\sigma$.

Proposition 4.4. (Cut elimination) The cut rule is redundant in HOL-$\lambda\sigma$.

Proof. Following the method developed in (Dowek and Werner, 1999) we construct a pre-model of the above rewrite system. A pre-model is a many-valued model whose truth values are reducibility candidates, i.e. sets of proof-terms. Hence we will first define proof-terms, then reducibility candidates and pre-models and at last construct a pre-model for HOL-$\lambda\sigma$.

Proof-terms are inductively defined as follows.

$$\begin{array}{rl}\pi ::= & \alpha \\ & \mid \lambda\alpha\ \pi \mid (\pi\ \pi') \\ & \mid (\pi, \pi') \mid fst(\pi) \mid snd(\pi) \\ & \mid i(\pi) \mid j(\pi) \mid (\delta\ \pi_1\ \alpha\pi_2\ \beta\pi_3) \\ & \mid (botelim\ \pi) \\ & \mid \lambda x\ \pi \mid (\pi\ t) \\ & \mid (t, \pi) \mid (exelim\ \pi\ x\alpha\pi')\end{array}$$

Each proof-term construction corresponds to a natural deduction rule: terms of the form α express proofs built with the axiom rule, terms of the form $\lambda\alpha\ \pi$ and $(\pi\ \pi')$ express proofs built with the introduction and elimination rules of the implication, terms of the form (π, π') and $fst(\pi)$, $snd(\pi)$ express proofs built with the introduction and elimination rules of the conjunction, terms of the form $i(\pi), j(\pi)$ and $(\delta\ \pi_1\ \alpha\pi_2\ \beta\pi_3)$ express proofs built with the introduction and elimination rules of the disjunction, terms of the form $(botelim\ \pi)$ express proofs built with the elimination rule of the contradiction, terms of the form $\lambda x\ \pi$ and $(\pi\ t)$ express proofs built with the introduction and elimination rules of the universal quantifier and terms of the form (t, π) and $(exelim\ \pi\ x\alpha\pi')$ express proofs built with the introduction and elimination rules of the existential quantifier.

Reduction on these proof-terms is defined by the following rules that eliminate cuts step by step.

$$(\lambda\alpha\ \pi_1\ \pi_2) \rhd \{\pi_2/\alpha\}\pi_1$$

$$fst(\pi_1, \pi_2) \rhd \pi_1$$

$$snd(\pi_1, \pi_2) \rhd \pi_2$$

$$(\delta\ i(\pi_1), \alpha\pi_2, \beta\pi_3) \rhd \{\pi_1/\alpha\}\pi_2$$

$$(\delta\ j(\pi_1), \alpha\pi_2, \beta\pi_3) \rhd \{\pi_1/\beta\}\pi_3$$

$$(\lambda x\ \pi\ t) \rhd \{t/x\}\pi$$

$$(exelim\ (t, \pi_1)\ \alpha x\pi_2) \rhd \{t/x, \pi_1/\alpha\}\pi_2$$

$$(\delta\ \pi_1\ \alpha\pi_2\ \beta\pi_3) \rhd \pi_2$$

$$(\delta\ \pi_1\ \alpha\pi_2\ \beta\pi_3) \rhd \pi_3$$

$$(exelim\ \pi_1\ x\alpha\pi_2) \rhd \pi_2$$

We are now ready to define reducibility candidates. We recall that a proof-term is said to be *neutral* if it is a proof variable or an elimination (i.e. of the form $(\pi\ \pi')$, $fst(\pi)$, $snd(\pi)$, $(\delta\ \pi_1\ \alpha\pi_2\ \beta\pi_3)$, $(botelim\ \pi)$, $(\pi\ t)$, $(exelim\ \pi\ x\alpha\pi')$), but not an introduction. A set R of proof-terms is a *reducibility candidate* if

— if $\pi \in R$, then π is strongly normalizable,
— if $\pi \in R$ and $\pi \rhd \pi'$ then $\pi' \in R$,
— if π is neutral and if for every π' such that $\pi \rhd^1 \pi'$, $\pi' \in R$ then $\pi \in R$.

We write $\mathcal{C}$ for the set of all reducibility candidates.

A *pre-model* for a language $\mathcal{L}$ is given by:

— for each sort T a set $\mathcal{M}_T$,
— for each function symbol f (of rank $(T_1, \ldots, T_n, U)$) a function $\overline{f}$ of $\mathcal{M}_U^{\mathcal{M}_{T_1} \times \ldots \times \mathcal{M}_{T_n}}$,
— for each predicate symbol P (of rank $(T_1, \ldots, T_n)$) a function $\overline{P}$ of $\mathcal{C}^{\mathcal{M}_{T_1} \times \ldots \times \mathcal{M}_{T_n}}$.

Let $\mathcal{M}$ be a pre-model, t be a term and φ an assignment mapping all the free variables of t of sort T to elements of $\mathcal{M}_T$. We define the object $|t|_\varphi$ by induction over the structure of t.

— $|x|_\varphi = \varphi(x)$,
— $|f(t_1, \ldots, t_n)|_\varphi = \overline{f}(|t_1|_\varphi, \ldots, |t_n|_\varphi)$.

Let A be a proposition and φ an assignment mapping all the free variables of A of sort T to elements of $\mathcal{M}_T$. We define the set $|A|_\varphi$ of proofs by induction over the structure of A.

— A proof π is an element of $|P(t_1, \ldots, t_n)|_\varphi$ if it is an element of $\overline{P}(|t_1|_\varphi, \ldots, |t_n|_\varphi)$.
— A proof π is element of $|A \Rightarrow B|_\varphi$ if it is strongly normalizable and if when it reduces to a proof of the form $\lambda\alpha\pi_1$ then for every π' in $|A|_\varphi$, $\{\pi'/\alpha\}\pi_1$ is an element of $|B|_\varphi$.
— A proof π is an element of $|A \wedge B|_\varphi$ if it is strongly normalizable and if when it reduces to a proof of the form (π_1, π_2) then π_1 and π_2 are elements of $|A|_\varphi$ and $|B|_\varphi$.
— A proof π is an element of $|A \vee B|_\varphi$ if it is strongly normalizable and if when it reduces to a proof of the form $i(\pi_1)$ (resp. $j(\pi_2)$) then π_1 (resp. π_2) is an element of $|A|_\varphi$ (resp. $|B|_\varphi$).
— A proof π is an element of $|\bot|_\varphi$ if it is strongly normalizable.
— A proof π is an element of $|\forall x\ A|_\varphi$ if it is strongly normalizable and if when it reduces to a proof of the form $\lambda x\ \pi_1$ then for every term t of sort T (where T is the sort of x) and every element E of $\mathcal{M}_T$ $\{t/x\}\pi_1$ is an element of $|A|_{\varphi+(x,E)}$.
— A proof π is an element of $|\exists x\ A|_\varphi$ if it is strongly normalizable and if when it reduces to a proof of the form (t, π_1) then for every element E of $\mathcal{M}_T$ (where T is the sort of t) then π_1 is an element of $|A|_{\varphi+(x,E)}$.

A pre-model is said to be a *pre-model of a congruence* $\equiv$ if when $A \equiv B$ then for every assignment φ, $|A|_\varphi = |B|_\varphi$.

It is proved in (Dowek and Werner, 1999) that if a congruence $\equiv$ has a pre-model, then the cut rule is redundant in (intuitionistic) sequent calculus modulo $\equiv$.

To prove that the cut rule is redundant in sequent calculus modulo HOL-$\lambda\sigma$, we construct a pre-model of this theory. We let

— $\mathcal{M}_\iota = \{0\}$,
— $\mathcal{M}_o = \mathcal{C}$, i.e. the set of all reducibility candidates.
— $\mathcal{M}_{T\to U} = \mathcal{M}_U^{\mathcal{M}_T}$.
— $\mathcal{D}_{T_1,\dots,T_n\vdash U} = \mathcal{M}_U{}^{\mathcal{M}_{T_n}{}^{\cdots^{\mathcal{M}_{T_1}}}}$,
— $\mathcal{D}_{T_1,\dots,T_n\vdash U_1,\dots,U_p} = (\mathcal{M}_{U_1}\times\ldots\times\mathcal{M}_{U_p})^{\mathcal{M}_{T_n}{}^{\cdots^{\mathcal{M}_{T_1}}}}$.

Then we interpret the symbols of the language as follows.

— $\overline{1_A^\Gamma}$ is the function mapping $a_1,\dots,a_n$ to a_1.
— $\overline{\alpha^\Gamma_{A\to B,A}}$ is the function mapping a and b to the function mapping $c_1,\dots,c_n$ to $a(c_1,\dots,c_n)(b(c_1,\dots,c_n))$.
— $\overline{\lambda^\Gamma_{A,B}}$ is the identity function.
— $\overline{[]_A^{\Gamma,\Gamma'}}$ is the function mapping a and b to $b\circ a$.
— $\overline{id^\Gamma}$ is the identity function.
— $\overline{\uparrow^\Gamma_A}$ is the function mapping $a_1,\dots,a_n$ to $(a_2,\dots,a_n)$
— $\overline{\cdot_A^{\Gamma,\Gamma'}}$ is the function mapping a, b to the function mapping $c_1,\dots,c_n$ to $(a(c_1,\dots,c_n),b(c_1,\dots,c_n))$.
— $\overline{\circ^{\Gamma,\Gamma',\Gamma''}}$ is the function mapping a and b to $b\circ a$.
— $\overline{\dot\Rightarrow}$ is the function mapping a and b to the set of proofs π such that π is strongly normalizable and when π reduces to a proof of the form $\lambda\alpha\pi_1$ then for every π' in a, $\{\pi'/\alpha\}\pi_1$ is an element of b.
— $\overline{\dot\wedge}$ is the function mapping a and b to the set of proofs π such that π is strongly normalizable and when π reduces to a proof of the form (π_1,π_2) then π_1 is an element of a and π_2 is an element of b.
— $\overline{\dot\vee}$ is the function mapping a and b to the set of proofs π such that π is strongly normalizable and when π reduces to a proof of the form $i(\pi_1)$ then π_1 is an element of a and when π reduces to a proof of the form $j(\pi_2)$ then π_2 is an element of b.
— $\overline{\dot\neg}$ is the function mapping a to the set of proofs π such that π is strongly normalizable and when π reduces to a proof of the form $\lambda\alpha\pi_1$ then for every π' in a, $\{\pi'/\alpha\}\pi_1$ is strongly normalizable.
— $\overline{\dot\bot}$ if the set of strongly normalizable proofs.
— $\overline{\dot\forall_T}$ is the function mapping a to the set of proofs π such that π is strongly normalizable and when π reduces to a proof of the form $\lambda x\ \pi_1$ then for every term t of sort T (where T is the sort of x) and every element E of $\mathcal{M}_T$ $\{t/x\}\pi_1$ is an element of $(a\ E)$,
— $\overline{\dot\exists_T}$ is the function mapping a to the set of proofs π such that π is strongly normalizable and when π reduces to a proof of the form (t,π_1) then for every element E of $\mathcal{M}_T$ (where T is the sort of x) $\{t/x\}\pi_1$ is an element of $(a\ E)$.
— $\overline{\varepsilon}$ is the identity function.

We check that the rules of $\lambda\sigma\mathcal{L}$ are valid in this pre-model, hence $\lambda\sigma\mathcal{L}$ has a pre-model and the cut rule is redundant in sequent calculus modulo $\lambda\sigma\mathcal{L}$.

Following the technique introduced in (Dowek and Werner, 1999) we can lift the cut elimination result to the classical sequent calculus modulo $\lambda\sigma\mathcal{L}$. □

5. Embedding HOL-λ into HOL-λσ

We now want to prove that HOL-$\lambda\sigma$ is intentionally equivalent to the usual presentation of higher-order logic HOL-λ. First, we have to translate a λ-term, say a, onto a $\lambda\sigma$-term. To motivate the way the translation is done, we recall, following (Dowek et al., 1995), that bound variables of a serve for reduction but the free variables serve only to be instantiated during a search proof process. In other words, the term a can be seen as a term with holes (called a context in (Barendregt, 1984)) filled by its free variables. So, bound variables are translated on de Bruijn numbers, letting the $\lambda\sigma$-rules doing reduction. But, free variables must be translated onto variables to remain instantiable so the translation of a is an open term of the first-order theory. However, as explained in (Dowek et al., 1995), as capture has to be avoided, first-order substitution alone cannot perform the instantiations. We need to use a translation, called *pre-cooking*, which translates free variables on variables of the first-order theory, relocated by an appropriate $[\uparrow^n]$ operator according to the form of the context. Then, instantiating correctly free variables by a simple first-order substitution is recovered. Let us recall how pre-cooking is defined.

To each variable x of type T in the λ-calculus, we associate the variable x of sort $\vdash T$ in $\lambda\sigma$-calculus.

Definition 5.1. Let a be a λ-term. The pre-cooking of a is the $\lambda\sigma$-term defined by $a_F = F(a, [\,])$ where $F(a, l)$ is defined using the list of variables l ($[\,]$ being the empty list) by:

— $F((\lambda x.a), l) = \lambda(F(a, x.l))$,
— $F((a\ b), l) = F(a, l)F(b, l)$,
— $F(x, l) = 1[\uparrow^{k-1}]$, if x is the k-th variable of l
— $F(x, l) = x[\uparrow^n]$ where n is the length of l if x is a variable not occurring in l or a constant.

We can now state our main theorem:

Theorem 5.1. If $p_1, \ldots, p_n, q_1, \ldots, q_m$ are propositions in HOL-λ then $p_1, \ldots, p_n \vdash q_1, \ldots, q_m$ is provable in HOL-λ iff $\varepsilon(p_{1F}), \ldots, \varepsilon(p_{nF}) \vdash \varepsilon(q_{1F}), \ldots, \varepsilon(q_{nF})$ is provable in HOL-$\lambda\sigma$.

The proof of this result relies on the next propositions. The first one recalls that pre-cooking is an homomorphism compatible with the respective substitution mechanisms and equalities:

Proposition 5.1. (Dowek et al., 1995)

— If t has the type T then t_F has the sort $\vdash T$,

— $(\{a/x\}b)_F = \{x \mapsto a_F\}b_F$, where $\{x \mapsto a\}b$ denotes the first-order substitution of the (first-order) variable x by the term a in the term b (while $\{a/x\}b$ denotes the capture avoiding substitution of the λ-calculus.)
— $a =_{\beta\eta} b$ in λ-calculus if and only if $a_F =_{\lambda\sigma} b_F$ in $\lambda\sigma$-calculus.

The purpose of the following definition and propositions is to characterize the image of the pre-cooking mapping.

Definition 5.2. A *F-term* is a $\lambda\sigma$-term containing only variables whose sort have empty contexts. A *F-proposition* is a proposition of the form $\varepsilon(P)$ where P is a F-term.

Proposition 5.2. If t is a $\lambda\sigma\mathcal{L}$-normal F-term well-typed in the empty context then there is a λ-term u such that $t = u_F$.

Proof. We prove by induction on the structure of t that if t is a $\lambda\sigma\mathcal{L}$-normal F-term well-typed in a context Γ, then there is a term u and a sequence l of variables of the same length than Γ such that $t = F(u, l)$.

The only interesting case is when $t = x[s]$. This term is well-typed in a context Γ of length n thus s has type $\Gamma \vdash$ and it is normal, thus $s = \uparrow^n$. □

Proposition 5.3. Let $\Gamma \vdash \Delta$ be a sequent containing only F-propositions. Then, if this sequent has a proof, it also has a proof where all propositions are F-propositions and all the witnesses F-terms.

Proof. By induction on the size of a cut free proof of $\Gamma \vdash \Delta$.

— If the last rule is an axiom then the result is obvious.
— If the last rule is a structural one then we apply the induction hypothesis to the subproofs.
— If the last rule is a left rule, we apply the induction hypothesis to the subproofs. The only non trivial case is the left rule of the universal quantifier. The proof has the form

$$\dfrac{\dfrac{\pi}{\Gamma, R \vdash \Delta}}{\Gamma, \varepsilon(q) \vdash \Delta}\,(x, P, t)\ \forall\text{-l}$$

Where $\varepsilon(q) \equiv_{\lambda\sigma\mathcal{L}} \forall x\ P$ and $R \equiv_{\lambda\sigma\mathcal{L}} \{t/x\}P$. Hence $q \equiv_{\lambda\sigma\mathcal{L}} (\dot{\forall}\ p)$, $P \equiv_{\lambda\sigma\mathcal{L}} \varepsilon(p\ x)$ and $R \equiv_{\lambda\sigma\mathcal{L}} \varepsilon(p\ t)$. Call σ the substitution mapping each variable x of t of sort $A_1, \ldots, A_n \vdash B$ to the term $x'[\uparrow^n]$ where x' is a fresh variable of sort $\vdash B$. By induction on the structure of π, the proof $\sigma\pi$ is a proof of $\Gamma, \sigma R \vdash \Delta$, i.e. $\Gamma, \varepsilon(p\ \sigma t) \vdash \Delta$. We apply the induction hypothesis to the proof $\sigma\pi$. Hence, there is a proof π' of $\Gamma, \varepsilon(p\ \sigma t) \vdash \Delta$ where all propositions are F-propositions and all the witnesses F-terms.
We build the proof

$$\dfrac{\dfrac{\pi'}{\Gamma, \varepsilon(p\ \sigma t) \vdash \Delta}}{\Gamma, \varepsilon(q) \vdash \Delta}\,(x, \varepsilon(p\ x), \sigma t)\ \forall\text{-l}$$

— If the last rule is a right rule, we apply the induction hypothesis to the subproofs.

The only non trivial case is the right rule of the existential quantifier. We proceed as for the left rule of the universal quantifier.

□

We can now give the proof of theorem 5.1.

Proof. The direct sense is an easy induction on the structure of the proof in HOL-λ. As an example, we detail the case: the last rule of the proof is the left rule of the universal quantifier. The proof has the form

$$\frac{\dfrac{\pi}{\Gamma, q \vdash \Delta}}{\Gamma, p \vdash \Delta}\,(r,t)\ \dot{\forall}\text{-l}$$

where $p = (\dot{\forall}\ r)$ and $q = (r\ t)\downarrow$. Then $\varepsilon(p_F) = \varepsilon(\dot{\forall}\ r_F) \equiv_{\lambda\sigma\mathcal{L}} \forall x\ \varepsilon(r_F\ x)$. By induction hypothesis, there is a proof π' of the sequent $\Gamma_F, \varepsilon(q_F) \vdash \Delta_F$. We build the proof

$$\frac{\dfrac{\pi'}{\Gamma_F, \varepsilon(q_F) \vdash \Delta_F}}{\Gamma_F, \varepsilon(p_F) \vdash \Delta_F}\,(x, \varepsilon(r_F\ x), t_F)\ \forall\text{-l}$$

Conversely, by the proposition 5.3, we can build a proof of $\varepsilon(p_{1F}), \ldots, \varepsilon(p_{nF}) \vdash \varepsilon(q_{1F}), \ldots, \varepsilon(q_{mF})$ where all the propositions are F-propositions and all the witnesses F-terms. By induction on the structure of this proof we can build a proof of $p_1, \ldots, p_n \vdash q_1, \ldots, q_m$ in HOL-λ. As an example, we give the case of the left rule of the universal quantifier. The proof has the form

$$\frac{\dfrac{\pi}{\Gamma_F, \varepsilon(q_F) \vdash \Delta_F}}{\Gamma_F, \varepsilon(p_F) \vdash \Delta_F}\,(x, \varepsilon(r_F), t_F)\ \forall\text{-l}$$

where $\varepsilon(p_F) = \forall x\ \varepsilon(r_F)$ and $q_F \equiv_{\lambda\sigma\mathcal{L}} \{x \mapsto t_F\} r_F$. Hence $p_F \equiv_{\lambda\sigma\mathcal{L}} (\dot{\forall}\ s_F)$ and $r_F \equiv (s_F\ x)$ and $q_F \equiv_{\lambda\sigma\mathcal{L}} (s_F\ t_F) = (s\ t)_F$. By induction hypothesis, there exists a proof π' in HOL-λ of $\Gamma, q \vdash \Delta$. We build the proof

$$\frac{\dfrac{\pi'}{\Gamma, q \vdash \Delta}}{\Gamma, p \vdash \Delta}\,(s,t)\ \forall\text{-l}$$

□

6. Skolemization in HOL-λσ

Skolemization in higher-order logic is known to be more complicated than in first-order logic. Indeed, the naive skolemization rule in higher-order logic allows to transform some unprovable formulations of the axiom of choice into provable propositions. Thus the naive skolemization rule has to be restricted in such a way that skolemizing a proposition of the form

$$\forall x_1 \ldots \forall x_n \exists y\ (P\ x_1\ \ldots\ x_n\ y)$$

introduces a skolem symbol f^n that can only be used in a substitution only if it is applied to at least n terms such that their free variables are not bound above in the term. For instance the term $\lambda y\ (f^1\ x\ y)$ can be used in a substitution, while the terms f^1, $(F\ f^1)$ and $\lambda x\ (f^1\ x\ y)$ cannot (Miller's conditions) (Miller, 1983; Miller, 1987).

A further motivation for expressing higher-order logic as a first-order theory is to avoid this cumbersome rule by reusing the usual first-order skolemization rule. We show below that when we apply the first-order skolemization rule to HOL-C we get conditions on Skolem symbols that are variants to Miller's conditions. In HOL-$\lambda\sigma$ we get exactly Miller's conditions.

6.1. *Miller's conditions in HOL-λ*

The naive treatment of skolemization in higher-order logic, that skolemizes

$$\forall x\ \exists y\ (P\ x\ y)$$

as

$$\forall x\ (P\ x\ (f\ x))$$

introduces a constant f of type $T \to U$ (where T is the type of x and U that of y). But this skolemization rule is unsound. Indeed, the axiom of choice

$$\forall x\ \exists y\ (P\ x\ y) \Rightarrow \exists g\ \forall x\ (P\ x\ (g\ x))$$

is not provable in type theory (Andrews, 1972). Thus from the proposition

$$\forall x\ \exists y\ (P\ x\ y)$$

we cannot deduce

$$\exists g\ \forall x\ (P\ x\ (g\ x))$$

while naively skolemizing it yields

$$\forall x\ (P\ x\ (f\ x))$$

from which we can obviously deduce

$$\exists g\ \forall x\ (P\ x\ (g\ x)).$$

Miller (Miller, 1983; Miller, 1987) has proposed an alternative skolemization rule that skolemizes a proposition of the form

$$\forall x_1\ \forall x_2\ \ldots\ \forall x_n\ \exists y\ (P\ x_1\ x_2\ \ldots\ x_n\ y)$$

into

$$\forall x_1\ \forall x_2\ \ldots\ \forall x_n\ (P\ x_1\ x_2\ \ldots\ x_n\ (f^n\ x_1\ x_2\ \ldots\ x_n)).$$

Two conditions are added to the terms substituted for variables:

— the symbol f^n can be used only when applied to at least n arguments (e.g. $(f^1\ x)$ can be used in a substitution, but f^1 alone cannot).

— the variables free in the necessary arguments cannot be bound by a λ above in the term (e.g. $\lambda x\ (f^1\ y)$ can be used in a substitution, but $\lambda x\ (f^1\ x)$ cannot).

Remark 6.1. As it is usual in higher-order logic, $\forall x\ P$ is a notation for the term $\dot{\forall}\ (\lambda x\ P)$ where $\dot{\forall}$ is a constant. With such a convention, the skolemized proposition $\forall x\ (P\ x\ (f^1\ x))$ itself does not verify the second condition since x is bound by the external quantifier. However this does not rule out this proposition because Miller's conditions do not apply to all terms and propositions, but only to the terms substituted for variables.

6.2. *Combinators*

Skolem theorem applies to the first-order presentation of higher-order logic with combinators as it applies to any first-order theory. A proposition of the form

$$\forall x\ \exists y\ \varepsilon(P\ x\ y)$$

is skolemized as

$$\forall x\ \varepsilon(P\ x\ f(x))$$

but then f is not a constant of type $T \to U$ but a function symbol of rank $(T)U$. Hence f alone is not a term (as $+$ is not a term in first-order arithmetic) but $f(x)$ is. We get this way Miller's first condition. As there is no notion of binding, the second condition vanishes in this presentation.

6.3. *HOL-λσ*

Skolem theorem also applies to HOL-$\lambda\sigma$ as it applies to any first-order theory. A proposition of the form

$$\forall x\ \exists y\ \varepsilon(P\ x\ y)$$

is skolemized as

$$\forall x\ \varepsilon(P\ x\ f(x))$$

Again f is a unary function symbol and hence we get back Miller's first condition, but its rank is now $(\Gamma \vdash T)\Delta \vdash U$, i.e. it maps an argument of sort $\Gamma \vdash T$ into a term of sort $\Delta \vdash U$. The sort of the argument expresses exactly Miller's second condition as it restricts the free variables in this term.

When the context associated to all variables is empty, the proposition

$$\forall x\ \exists y\ \varepsilon(P\ x\ y)$$

is skolemized as

$$\forall x\ \varepsilon(P\ x\ f(x))$$

where f has rank $(\vdash T) \vdash U$ which requires the argument of f to be well typed in the empty context. For instance the λ-term $\lambda x\ (f^1\ x)$ which violates Miller's second condition, is expressed by the term $\lambda(f(1))$ that is not well typed, while the term $\lambda x\ (f^1\ y)$ that verifies Miller's second condition is expressed by the term $\lambda(f(y))$ that is well-typed.

Notice that the restriction is simpler in this case as it applies uniformly to all the terms of the language, not only to the terms substituted for variables (see remark 6.1). The proposition $A = \forall x\ \varepsilon(P\ x\ f(x))$ is well-formed since the variable x bound by the quantifier

$$\frac{\{A_1, \ldots, A_n, B_1, \ldots, B_m\}[E_1] \quad \{\neg C_1, \ldots, \neg C_p, D_1, \ldots, D_q\}[E_2]}{\{B_1, \ldots, B_m, D_1, \ldots, D_q\}[E_1 \cup E_2 \cup \{A_1 \ldots =^?_{\mathcal{E}} A_n =^?_{\mathcal{E}} C_1 \ldots =^?_{\mathcal{E}} C_p\}]} \textbf{ Ext. Res.}$$

$$\frac{C[E]}{c\ell(C[r]_p)[E \cup \{C_{|p} =^?_{\mathcal{E}} l\}]} \textbf{ Ext. Nar.} \quad \text{if } l \to r \in \mathcal{R} \text{ and } C_{|p} \text{ is not a variable}$$

Fig. 5. Extended narrowing and resolution (ENAR)

$\forall$ is a variable of first-order logic and not a de Bruijn index. But there exists no term t of type o such that $A \equiv_{\lambda\sigma\mathcal{L}} \varepsilon(t)$ since the only candidate would be $t = (\dot{\forall}\ \lambda(P\ 1\ f(1)))$ that is ill-formed.

7. Automated theorem proving in HOL-$\lambda\sigma$

We are now able to wrap-up the above ingredients to get a first-order presentation of higher-order resolution. To this end, as with any first-order theory modulo, we can use the ENAR method developed in (Dowek et al., 1998) to search proofs in HOL-$\lambda\sigma$.

7.1. *The ENAR method*

The ENAR method applies to congruences described by class rewrite systems, i.e. pairs composed of a rewrite system $\mathcal{R}$ rewriting atomic propositions to propositions and a set of equational axioms $\mathcal{E}$ equating terms with terms and defining a congruence denoted $=_{\mathcal{E}}$.

As compared to first-order resolution, the ENAR method first replaces unification by equational unification modulo $\mathcal{E}$. The unification problems are kept as constraints written $t =^?_{\mathcal{E}} u$ and a clause C constrained by a set of equations E is written $C[E]$. Hence, we construct refutations with the **Extended Resolution** rule presented in figure 5. Then, as $\mathcal{R}$ rewrites atomic propositions to non atomic ones, we need another rule that instantiates, rewrites and puts in clausal form the result using the operator $c\ell$. This rule is called **Extended Narrowing** by analogy with the narrowing rule of equational unification.

Theorem 7.1. (Dowek et al., 1998) Let $\mathcal{RE}$ be a confluent and weakly terminating class rewrite system such that the cut rule is redundant in sequent calculus modulo $\mathcal{RE}$. Then, the sequent

$$A_1, \ldots, A_n \vdash B_1, \ldots, B_m$$

is provable in sequent calculus modulo if and only if from the constrained clauses

$$c\ell(\{\{A_1\}, \ldots, \{A_n\}, \{\neg B_1\}, \ldots, \{\neg B_m\}\})[\emptyset]$$

we can derive the empty clause constrained by a $\mathcal{E}$-unifiable set of equations.

7.2. *Applying ENAR to HOL-λσ*

In (Dowek et al., 1998), we have applied ENAR to a first-order presentation of higher-order logic using combinators. The system $\mathcal{E}$ contains the conversion rules of combinators while the system $\mathcal{R}$ contains the rules relating the connectors and quantifiers with their replication at the term level. We have shown that the **Extended Narrowing** rule specializes to the **Splitting** rule of higher-order resolution (Huet, 1972; Huet, 1973). Unfortunately equational unification modulo the conversion axioms of combinators is not higher-order unification.

If we apply this method to HOL-$\lambda\sigma$, we obtain another proof search method for higher-order logic. As shown in the previous sections, HOL-$\lambda\sigma$ fulfills the hypotheses of theorem 7.1, so this method is complete. The **Extended Narrowing** rule still specializes to the **Splitting** rule of higher-order resolution, but the unification required is the unification modulo the system $\lambda\sigma$ that we have shown to be equivalent to higher-order unification in (Dowek et al., 1995). Thus, the method obtained this way simulates higher-order resolution step by step.

Conclusion

In this paper we have given a first-order presentation of higher-order logic. This presentation is intentionally equivalent to the presentation of higher-order logic based on λ-calculus. Applying the Extended Narrowing and Resolution method to this theory gives exactly higher-order resolution. Hence we show this way that expressing higher-order logic as a first-order theory and applying a first-order proof search method is at least as efficient as a direct implementation, provided we take the right first-order expression of higher-order logic and the right proof search method.

Expressing higher-order resolution in a first-order framework allows to clarify its features: higher-order unification, the splitting rule and higher-order skolemization. Higher-order unification is equational unification in an appropriate theory. The splitting rule is an instance of the extended narrowing rule introduced in (Dowek et al., 1998), it is needed because the rewrite system of higher-order logic transforms atomic propositions into non atomic ones. The higher-order skolemization rule is an instance of the first-order one. Its scoping particularities are consequences of the sort system of higher-order logic.

Since we stay in a first-order setting, we can first reuse optimizations of first-order theorem proving such as redundancy criteria and subsumption. Second, extending the method to equational higher-order resolution requires only to add more reduction rules to the rewrite system $\lambda\sigma\mathcal{L}$, then narrowing provides an equational higher-order unification algorithm (Kirchner and Ringeissen, 1997) and the proof search method is complete provided deduction modulo the extended theory verifies the cut elimination property.

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