Introducing Quantified Cuts in Logic with Equality

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Abstract. Cut-introduction is a technique for structuring and compressing formal proofs. In this paper we generalize our cut-introduction method for the introduction of quantified lemmas of the form ∀x.A (for quantifier-free A) to a method generating lemmas of the form ∀x₁...∀xₙ.A. Moreover, we extend the original method to predicate logic with equality. The new method was implemented and applied to the TSTP proof database. It is shown that the extension of the method to handle equality and quantifier-blocks leads to a substantial improvement of the old algorithm.

1 Introduction

Computer-generated proofs are typically analytic, i.e., they only contain logical material that also appears in the statement of the theorem. This is due to the fact that analytic proof systems have a considerably smaller search space which makes proof-search practically feasible. In the case of sequent calculus, proof-search procedures typically work on the cut-free fragment. But also resolution is essentially analytic as resolution proofs satisfy the subformula property of first-order logic. One interesting property of non-analytic proofs is their considerably smaller length. The exact difference depends on the logic (or theory) under consideration, but it is typically enormous. In (classical and intuitionistic) first-order logic there are proofs with cut of length n whose theorems have only cut-free proofs of length 2ⁿ (where 2₀ = 1 and 2ⁿ₊₁ = 2²ⁿ) (see [15] and [12]). The length of a proof plays an important role in many situations such as human readability, space requirements and time requirements for proof checking. For most of these situations general-purpose data compression methods cannot be used as the compressed representation is not a proof anymore. It is therefore of high practical interest to develop methods of proof transformation which produce non-analytic and hence potentially much shorter proofs.

Work on cut-introduction can be found at a number of different places in the literature. Closest to our work are other approaches which aim to abbreviate or structure a given input proof. In [19] an algorithm for the introduction of atomic cuts that is capable of exponential proof compression is presented. The method [5] for propositional logic is shown to never increase the size of
proofs more than polynomially. Another approach to the compression of first-order proofs by introduction of definitions for abbreviating terms is \[18\]. There is a large body of work on the generation of non-analytic formulas carried out by numerous researchers in various communities. Methods for lemma generation are of crucial importance in inductive theorem proving which frequently requires generalization \[1\], see e.g. \[10\] for a method in the context of rippling \[2\] which is based on failed proof attempts. In automated theory formation \[3,4\], an eager approach to lemma generation is adopted. This work has, for example, led to automated classification results of isomorphism classes \[13\] and isotopy classes \[14\] in finite algebra. See also \[11\] for an approach to inductive theory formation.

Methods of algorithmic cut-introduction, based on the inversion of Gentzen’s cut-elimination method, have been defined in \[8\] and \[7\]. The method in \[8\] works on a cut-free \(\text{LK}\)-proof \(\varphi\) of a prenex skolemized end-sequent \(S\) and consists of the following steps: (1) extraction of a set of terms \(T\) from \(\varphi\), (2) computation of a compressed representation of \(T\), (3) construction of the cut formula, (4) improvement of the solution by computation smaller cut-formulas, and (5) construction of an \(\text{LK}\)-proof with the universal cut formula obtained in (4) and instantiation of the quantifiers with the terms obtained in (2). It has been shown in \[8\] that the method is capable of compressing cut-free proofs quadratically. The paper \[7\] generalized the method to the introduction of arbitrarily many universal cut formulas, where the steps defined above are roughly the same, though the improvement of the solution (step 4) and the final construction of the proof with cuts (step 5) are much more difficult. The method of introducing arbitrarily many universal cuts in \[7\] leads even to an exponential compression of proof length. Still the methods described above were mainly designed for a theoretical analysis of the cut-introduction problem rather than for practical applications. In particular, they lacked efficient handling of equality (as they were defined for predicate logic without equality) and the introduction of several universal quantifiers in cut formulas (all cut formulas constructed in \[7\] are of the form \(\forall x.A\) for a single variable \(x\) and a quantifier-free formula \(A\)).

In this paper we generalize our cut-introduction method to predicate logic with equality and to the construction of a (single) quantified cut containing blocks of universal quantifiers. The efficient compression of the terms (step 2) and the improvement of the solution (step 4) require new and non-trivial techniques. Moreover, we applied the new method in large-scale experiments to proofs generated by prover9 on the TPTP library. This empirical evaluation demonstrates the feasibility of our method on realistic examples.

## 2 Proofs and Herbrand Sequents

Throughout this paper we consider predicate logic with equality. For practical reasons equality will not be axiomatized but handled via substitution rules. We extend the sequent calculus \(\text{LK}\) to the calculus \(\text{LK}_e\) by allowing sequents of the form \(\rightarrow t = t\) as initial sequents and adding the following rules:
$\Gamma \rightarrow \Delta, s = t \quad A[s], \Pi \rightarrow A$  
\hspace{1cm} \text{El1}  
$A[t], \Gamma, \Pi \rightarrow \Delta, A$  

$\Gamma \rightarrow \Delta, s = t \quad II \rightarrow A[s], A$  
\hspace{1cm} \text{Er1}  
$\Gamma, \Pi \rightarrow A[t], \Delta, A$  

$\Gamma \rightarrow \Delta, t = s \quad A[s], \Pi \rightarrow A$  
\hspace{1cm} \text{El2}  
$A[t], \Gamma, \Pi \rightarrow \Delta, A$  

$\Gamma \rightarrow \Delta, t = s \quad II \rightarrow A[s], A$  
\hspace{1cm} \text{Er2}  
$\Gamma, \Pi \rightarrow A[t], \Delta, A$  

$LK_e$ is sound and complete for predicate logic with equality.

For convenience we write a substitution $[x_1/t_1, \ldots, x_n/t_n]$ in the form $[\bar{x}/\bar{t}]$ for $\bar{x} = (x_1, \ldots, x_n)$ and $\bar{t} = (t_1, \ldots, t_n)$. A strong quantifier is a $\forall$ ($\exists$) quantifier with positive (negative) polarity. We restrict our investigations to end-sequents in prenex form without strong quantifiers.

**Definition 1.** A $\Sigma_1$-sequent is a sequent of the form

$$\forall x_1 \cdots \forall x_{k_1} F_1, \ldots, \forall x_1 \cdots \forall x_{k_p} F_p \rightarrow \exists x_1 \cdots \exists x_{k_{p+1}} F_{p+1}, \ldots, \exists x_1 \cdots \exists x_{k_q} F_q,$$

for quantifier free $F_i$.

Note that the restriction to $\Sigma_1$-sequents does not constitute a substantial restriction as one can transform every sequent into a validity-equivalent $\Sigma_1$-sequent by skolemisation and prenexing.

**Definition 2.** A sequent $S$ is called E-valid if it is valid in predicate logic with equality; $S$ is called a quasi-tautology if $S$ is quantifier-free and E-valid.

**Definition 3.** The length of a proof $\varphi$, denoted by $|\varphi|$, is defined as the number of inferences in $\varphi$. The quantifier-complexity of $\varphi$, written as $|\varphi|_q$, is the number of weak quantifier-block introductions in $\varphi$.

### 2.1 Extraction of terms

Herbrand sequents of a sequent $S$ are sequents consisting of instantiations of $S$ which are quasi-tautologies. The formal definition is:

**Definition 4.** Let $S$ be a $\Sigma_1$-sequent as in Definition 1 and let $H_i$ be a finite set of $k_i$-vectors of terms for every $i \in \{1, \ldots, q\}$. We define $F_i = \{F_i[\bar{x}/\bar{t}] \mid \bar{t} \in H_i\}$ if $k_i > 0$ and $F_i = \{F_i\}$ if $k_i = 0$. Let

$$S^* : F_1 \cup \cdots \cup F_p \rightarrow F_{p+1} \cup \cdots \cup F_q.$$  

If $S^*$ is a quasi-tautology then it is called a Herbrand sequent of $S$ and $H : (H_1, \ldots, H_q)$ is called a Herbrand structure of $S$. We define the size of $S^*$ as $\sum_{i=1}^q |H_i|$.

Note that, in the size of a Herbrand sequent, we only count the formulas obtained by instantiation.
Example 1. Consider the language containing a constant symbol \( a \), unary function symbols \( f, s \), a binary predicate symbol \( P \), and the sequent \( S \) defined below. We write \( f^n, s^n \) for \( n \)-fold iterations of \( f \) and \( s \) and omit parentheses around the argument of a unary symbol when convenient. Let

\[
S : P(f^4a, a), \forall x. f x = s^2 x, \forall xy(P(sx, y) \supset P(x, sy)) \rightarrow P(a, f^4 a)
\]

and \( H = (H_1, H_2, H_3, H_4) \) for

\[
\begin{align*}
H_1 & = \emptyset, \quad H_4 = \emptyset, \\
H_2 & = \{ a, f a, f^2 a, f^3 a \}, \\
H_3 & = \{ (s^3 f^2 a, a), (s^2 f^2 a, sa), (s f^2 a, s^2 a), (f^2 a, s^3 a), (s^3 a, f^2 a), (s^2 a, s f^2 a), (sa, s^2 f^2 a), (a, s^3 f^2 a) \}.
\end{align*}
\]

Then

\[
\begin{align*}
F_1 & = \{ P(f^4 a, a) \}, \quad F_4 = \{ P(a, f^4 a) \}, \\
F_2 & = \{ fa = s^2 a, f^2 a = s^2 fa, f^3 a = s^2 f^2 a, f^4 a = s^2 f^3 a \}, \\
F_3 & = \{ P(s^3 f^2 a, sa), P(s^2 f^2 a, sa) \supset P(s^2 f^2 a, s^2 a), P(s^2 f^2 a, s^2 a) \supset P(s f^2 a, s^3 a), P(s^2 f^2 a, s^3 a) \supset P(f^2 a, s^3 a), P(s f^2 a, s^3 a) \supset P(s^3 a, f^2 a), P(s^3 a, f^2 a) \supset P(s^2 a, s^2 f^2 a), P(s^2 a, s^2 f^2 a) \supset P(s a, s^3 f^2 a), P(s a, s^3 f^2 a) \supset P(a, s^3 f^2 a) \}.
\end{align*}
\]

A Herbrand-sequent \( S^* \) corresponding to \( H \) is then \( F_1 \cup F_2 \cup F_3 \rightarrow F_4 \). Note that \( fa = s^2 a, f^2 a = s^2 fa, f^3 a = s^2 f^2 a, f^4 a = s^2 f^3 a, f^4 a = s^2 f^3 a \).

The size of \( S^* \) is 12. \( S^* \) is a quasi-tautology but not a tautology.

**Theorem 1 (mid-sequent theorem).** Let \( S \) be a \( \Sigma_1 \)-sequent and \( \pi \) a cut-free proof of \( S \). Then there is a Herbrand-sequent \( S^* \) of \( S \) s.t. \( |S^*| \leq |\pi|_q \).

**Proof.** This result is proven in [6] for \( \mathbf{LK} \), but the proof for \( \mathbf{LK}_\omega \) is basically the same. By permuting the inference rules, one obtains a proof \( \pi' \) from \( \pi \) which has an upper part containing only propositional inferences and the equality rules (which can be shifted upwards until they are applied to atoms only) and a lower part containing only quantifier inferences. The sequent between these parts is called *mid-sequent* and has the desired properties.

\( S^* \) can be obtained by tracing the introduction of quantifier-blocks in the proof, which for every formula \( Q \bar{x}, F_i \) in the sequent (where \( Q \in \{ \forall, \exists \} \) yields a set of term tuples \( H_i \), and then computing the sets of formulas \( F_j \).

The algorithm for introducing cuts described here relies on computing a compressed representation of the Herbrand structure, which is explained in Section 9.

Note, though, that the Herbrand structure \( (H_1, \ldots, H_q) \) is a list of sets of term tuples (i.e. each \( H_i \) is a set of tuples \( \bar{t} \) used to instantiate the formula \( F_i \)). In order to facilitate computation and representation, we will add to the language fresh function symbols \( f_1, \ldots, f_q \). Each \( f_i \) will be applied to the tuples of the set \( H_i \), therefore transforming a list of sets of tuples into a set of terms. In this new set, each term will have an \( f_k \) as its head symbol, that indicates to which formula the arguments of \( f_k \) belong.

Example 2. Using this new notation, the Herbrand structure \( H \) of the previous example is now represented as the set of terms:

\[
T : \{ f_2(a), f_2(fa), f_2(f^2 a), f_2(f^3 a), f_3(s^3 f^2 a, a), f_3(s^2 f^2 a, sa), \ldots, f_3(a, s^3 f^2 a) \}.
\]

Henceforth we will refer to the transformed Herbrand structure as the *term set* of a proof.
3 Computing a Decomposition

We shall now describe an algorithm for computing a compressed representation of a term set \( T \). Term sets will be represented by decompositions which are defined as follows:

**Definition 5.** Let \( T = \{t_1, \ldots, t_n\} \) be a set of ground terms. A decomposition \( D \) of \( T \) is a pair, written as \( U \circ_\alpha W \), where \( U \) is a set of terms containing the variables \( \alpha_1, \ldots, \alpha_m \), and \( W = \left\{ \bar{w}_1 = \left( \begin{array}{c} w_{1,1} \\ \vdots \\ w_{1,m} \end{array} \right), \ldots, \bar{w}_q = \left( \begin{array}{c} w_{q,1} \\ \vdots \\ w_{q,m} \end{array} \right) \right\} \) is a set of vectors of ground terms s.t. the following two conditions are satisfied:

1. For every entry \( W \Rightarrow \{ (u_1, T_1), \ldots, (u_n, T_n) \} \), \( \{ u_i \} \circ W \) is a decomposition of \( T_i \) (for \( 1 \leq i \leq n \)).
2. For every \( T' \subseteq T \), there is a pair \( W \Rightarrow U_\Delta \) in the \( \Delta \)-table s.t. \( (u, T') \in U_\Delta \).

We build the \( \Delta \)-table as follows: for every \( T' \subseteq T \), we compute \( \Delta_G(T') = (u, W) \) and insert \( (u, T') \) with the key \( W \) (if an entry \( W \Rightarrow U_\Delta \) already exists, we replace it with \( W \Rightarrow U_\Delta \cup \{ (u, T') \} \)). We then iterate over the \( \Delta \)-table and, for each entry \( W \Rightarrow U_\Delta \), try to find a subset \( \{ (u_{i_1}, T_{i_1}), \ldots, (u_{i_n}, T_{i_n}) \} \) of \( U_\Delta \) s.t. \( \{ u_{i_1}, \ldots, u_{i_n} \} \circ W \) is a decomposition of \( T \). This is called folding the \( \Delta \)-table.

**Theorem 2 (Soundness and completeness).** Let \( T \) be a term set. If \( U \circ W \) is extracted from folding the \( \Delta \)-table, then \( U \circ W \) is a decomposition of \( T \). Conversely, if there exists a decomposition \( U \circ W \) of \( T \) of size \( n \), folding the \( \Delta \)-table will return at least one decomposition of size \( n' \leq n \).

*Proof.* See Appendix.
insert a new entry into a decomposition of $T$. We find the entry $u,T \in U$ to which $\Delta$-vector $\Delta_U$ has been computed for all subsets, consider for instance the subset $T' = \{ f_3(s^2f^2a,a), f_3(s^2f^2a,sa) \} \subseteq T$:

$$\Delta_U(f_3(s^2f^2a,a), f_3(s^2f^2a,sa)) = (f_3(s^2f^2a,a), \begin{pmatrix} s^2f^2a \\ f^2a \end{pmatrix}, \begin{pmatrix} f^2a \\ sa \end{pmatrix}) = (u,W).$$

If the $\Delta$-table already has an entry $W \Rightarrow U_\Delta$, we add $(u,T')$ to $U_\Delta$. If not, we insert a new entry $W \Rightarrow \{(u,T')\}$. After $\Delta_U$ has been computed for all subsets, we iterate through it, looking for simple decompositions that can be composed into a decomposition of $T$. We find the entry

Algorithm 1 Generalized $\Delta$-vector $\Delta_G$

function $\Delta_G(t_1,\ldots,t_n)$; a list of terms
    return transposeW($\Delta_G(t_1,\ldots,t_n)$)
end function

function $\Delta_G(t_1,\ldots,t_n)$; a list of terms
    if $t_1 = t_2 = \ldots = t_n \wedge n \geq 0$ then \(\triangleright\) case 1: all terms identical
        return ($t_1,())$
    else if $t_i = f(t_{i1},\ldots,t_{in})$ for $1 \leq i \leq n$ then \(\triangleright\) case 2: recurse
        ($\bar{w}_1,\ldots,\bar{w}_m) \leftarrow \bigcup_{1 \leq j \leq m} \pi_2(\Delta_G(t_{j1},\ldots,t_{jn}))$
        $u_j \leftarrow \pi_1(\Delta_G(t_{j1},\ldots,t_{jn}))$ for all $j \in \{1,\ldots,m\}$
        return merge($f(u_1,\ldots,u_m),(\bar{w}_1,\ldots,\bar{w}_m)$) \(\triangleright\) case 3: introduce new $\alpha$
    else
        return $(\alpha\text{FRESH},(t_1,\ldots,t_n))$
end if
\[ W \Rightarrow U^1_\Delta \cup U^2_\Delta \]
\[ U^1_\Delta = \{(f_2(a_1), \{t_1, t_2\}), (f_2(f_3), \{t_2, t_4\}), (f_2(a_2), \{t_1, t_3\}) (f_2(f_3), \{t_2, t_4\}) \]
\[ U^2_\Delta = \{(f_3(s^3 a_1, a_2), \{t_5, t_6\}), (f_3(s^2 a_1, a_2), \{t_6, t_{10}\}), (f_3(s^2 a_1, s^2 a_2), \{t_7, t_{11}\}), (f_3(a_1, s^3 a_2), \{t_8, t_{12}\}) \]
\[ W = \left\{ \left(\frac{f^2 a}{a}, \frac{a}{f^2 a} \right) \right\}. \]

and can see that \( U^1_\Delta \circ W = T_2 \) and \( U^2_\Delta \circ W = T_3 \). Therefore, \( (U^1_\Delta \cup U^2_\Delta) \circ W \) is a decomposition of \( T \). We then translate this decompositions \( T \) back into a decomposition of \( H \) by removing the function symbols \( f_2 \) and \( f_3 \) from \( U^1_\Delta \cup U^2_\Delta \):

\[ U = (U_1, U_2), \]
\[ U_1 = \{a_1, f a_1, a_2, f a_2\}, \]
\[ U_2 = \{(s^3 a_1, a_2), (s^2 a_1, s a_2), (s a_1, s^2 a_2), (a_1, s^3 a_2)\}, \]
\[ W = \left\{ \left(\frac{f^2 a}{a}, \frac{a}{f^2 a} \right) \right\}. \]

4 Computing a Cut-Formula

After having computed a decomposition as described in Section 3, the next step consists in computing a cut-formula based on that decomposition. A decomposition \( D \) specifies the instances of quantifier blocks in a proof with a \( \forall \)-cut, but does not contain information about the propositional structure of the cut formula to be constructed. The problem to find the appropriate propositional structure is reflected in the following definition.

**Definition 7.** Let \( S \) be a \( \Sigma_1 \)-sequent and \( F_i, k_i \) as in Definition \( 7 \), \( H \) be a Herbrand structure for \( S \), and \( D: U \circ W \) a decomposition of \( H \) with \( V(D) = \{a_1, \ldots, a_n\} \). Let \( U = (U_1, \ldots, U_q) \) and \( W = \{\bar{w}_1, \ldots, \bar{w}_k\} \), where the \( \bar{w}_i \) are \( n \)-vectors of terms not containing variables in \( V(D) \), and \( F'_i = \{F_i[x]\bar{\tilde{t}} | \tilde{t} \in U_i\} \) for \( k_i > 0 \) and \( F'_i = \{F_i\} \) for \( k_i = 0 \). Furthermore let \( X \) be an \( n \)-place predicate variable. Then the sequent

\[
S^\sim : X\bar{\tilde{a}} \supset \bigwedge_{i=1}^k X\bar{w}_i, F'_1, \ldots, F'_p \rightarrow F'_{p+1}, \ldots, F'_{q'}, \]

is called a schematic extended Herbrand sequent of \( S \) w.r.t. \( D \). The size of \( S^\sim \), denoted by \( |S^\sim| \), is defined as \( k + \sum_{i=1}^q |U_i| \).

**Definition 8.** Let \( S^\sim \) be a schematic extended Herbrand sequent of \( S \) w.r.t. a decomposition \( D \) as in Definition 7 and \( A \) be a formula with \( V(A) \subseteq \{a_1, \ldots, a_n\} \). Then the second-order substitution \( \sigma; [X\bar{\tilde{a}}.A] \) is a solution of \( S^\sim \) if \( S^\sim \sigma \) is a quasi-tautology; in this case \( S^\sim \sigma \) is called an extended Herbrand sequent. The size of \( S^\sim \sigma \) is defined as \( |S^\sim| \).
Theorem 6 in Section 4.3 shows that, from a solution of a schematic extended Herbrand sequent $S^\sim$ of $S$, we can define a proof $\psi$ of $S$ with a $\forall$-cut and $|\psi|_q = |S^\sim|$. The question remains whether every schematic extended Herbrand sequent is solvable. We show below that this is indeed the case.

Let $S^\sim$ as in Definition 7. We define

$$F[l] = \bigwedge_{i=1}^{p} F'_i \quad \text{and} \quad F[r] = \bigvee_{i=p+1}^{q} F'_i.$$  

**Definition 9.** Let $S^\sim$ be a schematic extended Herbrand sequent of $S$ as in Definition 7. We define the canonical formula $C(S^\sim)$ of $S^\sim$ as $F[l] \land \neg F[r]$. The substitution $[X\backslash \lambda \bar{\alpha} C(S^\sim)]$ is called the canonical substitution of $(S,S^\sim)$.

**Theorem 3.** Let $S$ be a $\Sigma_1$-sequent, and $S^\sim$ be a schematic extended Herbrand sequent of $S$. Then the canonical substitution is a solution of $S^\sim$.

**Proof.** Let $S^\sim$ be a schematic extended Herbrand sequent as in Definition 7 and $C(S^\sim)$ be the canonical formula of $S^\sim$. We have to prove that

$$S_1 : C(S^\sim)(\bar{\alpha}) \supset \bigwedge_{i=1}^{k} C(S^\sim)(\bar{w}_i), \quad F[l] \rightarrow F[r]$$

is a quasi-tautology. But, by definition of $C(S^\sim)$, $S_1$ is equivalent to

$$S_2 : (F[l] \land \neg F[r]) \supset \bigwedge_{i=1}^{k} (F[l] \land \neg F[r])(\bar{w}_i), \quad (F[l] \land \neg F[r]) \rightarrow .$$

Clearly $S_2$ is a quasi-tautology if the sequent $S_3$, defined as

$$S_3 : \bigwedge_{i=1}^{k} (F[l] \land \neg F[r])(\bar{w}_i) \rightarrow$$

is a quasi-tautology. But, by $D = U \circ W$ being a decomposition of $H$, $S_3$ is logically equivalent to the Herbrand sequent $S^*$ defined over $H$, which (by definition) is a quasi-tautology.

**Example 4.** Let

$$S : P(f^4 a, a), \forall x.f x = s^2 x, \forall x y (P(s x, y) \supset P(x, s y)) \rightarrow P(a, f^4 a)$$

like in Example 1 and $D$ be the decomposition $U \circ W$ of $H$ constructed in Example 3. We have

$$U = (U_1, U_2),$$

$$U_1 = \{\alpha_1, f \alpha_1, \alpha_2, f \alpha_2\},$$

$$U_2 = \{(s^3 \alpha_1, \alpha_2), (s^2 \alpha_1, s \alpha_2), (s \alpha_1, s^2 \alpha_2), (\alpha_1, s^3 \alpha_2)\},$$

$$W = \left\{ \left( \begin{array}{c} f^2 a \\ a \end{array} \right), \left( \begin{array}{c} a \\ f^2 a \end{array} \right) \right\}.$$
The corresponding schematic extended Herbrand sequent \( S^\sim \) is

\[
X(\alpha_1, \alpha_2) \supset (X(f^2 a, a) \land X(a, f^2 a)),
\]

\( f_{\alpha_1} = s^2 \alpha_1, \ f^2 \alpha_1 = s^2 f_{\alpha_1}, \ f_{\alpha_2} = s^2 \alpha_2, \ f^2 \alpha_2 = s^2 f_{\alpha_2} \),

\[
P(s^3 \alpha_1, \alpha_2) \supset P(s^3 \alpha_1, so_2), \ P(s^3 \alpha_1, so_2) \supset P(s^2 \alpha_1, s^2 \alpha_2), \ P(s^2 \alpha_1, s^2 \alpha_2) \supset P(so_1, s^3 \alpha_2),
\]

\[
P(so_1, s^3 \alpha_2) \supset P(\alpha_1, s^3 \alpha_2), \ P(f^4 a, a) \rightarrow P(a, f^4 a).
\]

Its canonical formula \( C(S^\sim) \) which we write as \( A(\alpha_1, \alpha_2) \) is

\[
\bigwedge_{i=1}^3 (f_{\alpha_i} = s^2 \alpha_i \land f^2 \alpha_i = s^2 f_{\alpha_i}) \land \\
\bigwedge_{i=0}^3 (P(s^{4-i} \alpha_1, s^4 \alpha_2) \supset P(s^{4-i} \alpha_1, s^{i+1} \alpha_2)) \land P(f^4 a, a) \land \neg P(a, f^4 a).
\]

The canonical solution is \( [X \setminus \lambda \alpha_1 \alpha_2.A(\alpha_1, \alpha_2)] \) and the corresponding Herbrand sequent \( S' \) is like \( S^\sim \) with \( X(\alpha_1, \alpha_2) \supset (X(f^2 a, a) \land X(a, f^2 a)) \) replaced by \( A(\alpha_1, \alpha_2) \supset (A(f^2 a, a) \land A(a, f^2 a)) \). Note that \( |S'| = 10 \), while \( |S^\sim| = 12 \). So we obtained a compression of quantifier complexity.

### 4.1 Improving the solution

In the last section, we have shown that, given a decomposition \( D \) of the termset of a cut-free proof of a \( \Sigma_1 \)-sequent \( S \), there exists a canonical solution to the schematic extended Herbrand sequent induced by \( S, D \), which gives rise to a proof with a \( \forall \)-cut. The canonical solution need not be the best solution for a given purpose; indeed it is often not symbol-minimal, for example. Hence this section is devoted to describing an algorithm for finding better solutions. We will consider E-validity of quantifier-free formulas \( F \) containing free variables; by "\( F \) is E-valid" we mean to say "the universal closure of \( F \) is E-valid". Throughout this section, we consider a fixed \( \Sigma_1 \)-sequent \( S \) using the notation of Definition\(^{[1]} \) a fixed decomposition \( D = (U_1, \ldots, U_q) \circ W \), with \( W = \{ \bar{w}_i \mid 1 \leq i \leq k \} \), of a Herbrand structure \( H \) of \( S \), along with the schematic extended Herbrand sequent \( S^\sim \) induced by \( S, H, D \), using the notation of Definition\(^{[4]} \). We will abbreviate \( \mathcal{F}_p' \cup \cdots \cup \mathcal{F}_q' \) by \( \Gamma' \) and \( \mathcal{F}_{p+1}' \cup \cdots \cup \mathcal{F}_q' \) by \( \Delta \), and write "\( A \) is a solution" for "\( [X \setminus \lambda \bar{x}.A] \) is a solution for \( S^\sim \)" (note that we will consider the names \( \bar{x} \) fixed).

In this section, we will focus our attention on solutions in conjunctive normal form (CNF), which always exist since the solution property is semantic (if \( A \) is a solution and \( A \leftrightarrow B \) is E-valid, then \( B \) is a solution). A clause \( C \) is said to be \( \bar{x} \)-free if it contains no symbol from \( \bar{x} \).

The algorithm we will present will involve generating E-consequences of formulas. Although in principle an abstract analysis of our algorithm based on a notion of E-consequence generator can be performed, we have chosen, for lack of space, to present only the concrete E-consequence generator used in our implementation.

We now present this E-consequence generator, which is based on forgetful reasoning. Let \( C_1, C_2 \) be two clauses, then denote the set of propositional resolvents of \( C_1, C_2 \) by \( \text{res}(C_1, C_2) \) and the set of clauses that can be obtained from
$C_1, C_2$ by ground paramodulation by $\text{para}(C_1, C_2)$. Letting $F$ be a formula with CNF $\{C_i\}_{i \in I}$ we define

$$F(F) = \{ C \land \bigwedge_{i \in I \setminus \{j, k\}} C_i \mid C \in \text{res}(C_j, C_k) \cup \text{para}(C_j, C_k) \}. $$

Using $F$, we can now present Algorithm 2 the solution-finding algorithm $\text{SF}_F$. It prunes a solution $A$ of $\bar{x}$-free clauses, then recurses upon those consequences of the pruned $A$ generated by $F$ which pass a certain E-validity check, finally returning a set of formulas (which will all be solutions). We have the following result, which is derived essentially from the algebraic structure of the solution space which is sketched in the following section.

**Theorem 4 (Soundness & Termination).** Let $A$ be any solution in CNF. Then $\text{SF}_F$ terminates on $A$ and, for all $B \in \text{SF}_F(A)$, $B$ is a solution.

**Example 5.** Consider the canonical formula $C(S\sim)$ of Example 4. Then $\text{SF}_F$ generates the CNF

$$F(\alpha_1, \alpha_2) : f^2 \alpha_1 = s^4 \alpha_1 \land f^2 \alpha_2 = s^4 \alpha_2 \land (\neg P(s^4 \alpha_1, \alpha_2) \lor P(\alpha_1, s^4 \alpha_2))$$

for the CNF of $C(S\sim)$ by applying paramodulation twice to equational atoms and resolution thrice to the clauses corresponding to the implications between the $P$-atoms. It can be checked that $\lambda \alpha_1 \alpha_2. F(\alpha_1, \alpha_2)$ is a solution for $S\sim$ which is smaller than the canonical solution.

### 4.2 The solution space

This section is dedicated to describing the space of solutions. The following result summarizes the algebraic properties of the solution space which are exploited in $\text{SF}_F$; in particular these properties allow one to prove the correctness of $\text{SF}_F$ in the sense of Theorem 4. By $\mathcal{F}$ we denote the set of propositional formulas built from the atoms of the canonical solution of $S\sim$. 
Theorem 5. Define the equivalence relation $\sim$ on $\mathcal{F}$ by setting $F \sim G$ iff $F \leftrightarrow G$ is $E$-valid (i.e. $\mathcal{F}/\sim$ is the Lindenbaum-Tarski algebra of $\mathcal{F}$ w.r.t. the theory of equality). Let $\mathcal{S} \subseteq \mathcal{F}$ be the set of solutions in that signature. Then $B = (\mathcal{F}/\sim, \land, \lor, \neg, \bot, \top)$ is a Boolean algebra, $(\mathcal{S}/\sim, \land)$ is a convex subalgebra of the meet-semilattice reduct $(\mathcal{F}/\sim, \land)$ of $B$, and the canonical solution $C$ is the least element of $(\mathcal{S}/\sim, \land)$.

![Figure 1](image)

Figure 1 visualizes this result for a language with 4 atoms. The algorithm $SF$ can be seen as searching upwards through the solution semilattice, starting at the bottom element $C$. For any (representative of an) element $F$ of this semilattice, $F(F)$ contains formulas which either correspond to $F$, or an element above $F$. Whenever we reach an element $G$ that is not a solution, we know that no other element above $G$ is a solution since the solution space is convex.

As indicated above, the algorithm $SF$ can in principle be stated using $E$-consequence generators different from $F$, and one can formulate general properties of such generators from which important results, such as completeness, can be derived. Furthermore, it will be possible to apply methods from automated theorem proving in presence of equality to the development of practically useful $E$-consequence generators. We leave such further theoretical and empirical investigation of “good” $E$-consequence generators for future work.

4.3 Proof with cut

Theorem 6. Let $S^\sim$ be an extended Herbrand sequent of a $\Sigma_1$-sequent $S$. Then $S$ has a proof with a $\forall$-cut s.t. $|\varphi|_q = |S^\sim|$.

Proof. As in [7]. Note that the quantifier-blocks in the cut and the equality rules do not change the measured number of weak quantifier-block introductions.
analyzed in the paper above. The main steps in the proof are the following ones: let \( S' \) be an extended Herbrand sequent obtained by the solution \([X \backslash \lambda \bar{\alpha}, A]\). Then a proof with cut formula \( \forall \bar{x}.A[\bar{\alpha} \backslash \bar{x}] \) can be constructed where the quantifier substitution blocks for the cut formula on the right-hand-side are \([\bar{x} \backslash \bar{w}]\) for \( \bar{w} \in W \) while the cut formula on the left-hand-side gets the substitution \([\bar{x} \backslash \bar{\alpha}]\). The substitutions \([\bar{x}_i \backslash \bar{t}]\) for \( \bar{t} \in U_i \) are inserted to introduce the quantifiers of the formula \( F_i \) in the end-sequent.

**Example 6.** Let \( \Gamma = P(f^4 a, a), \forall x. f x = s^2 x, \forall x y (P(s x, y) \supset P(x, s y)) \) be the left-hand-side of \( S \). Then, to the canonical solution corresponds an \( \text{LK} \)-proof \( \psi \) of the form

\[
\frac{(\psi_1)}{\Gamma \rightarrow P(a, f^4 a), A(\alpha_1, \alpha_2)} \quad \frac{(\psi_2)}{\Gamma, A(f^2 a, a), A(a, f^2 a) \rightarrow P(a, f^4 a)}
\]

\[
\frac{\forall \bar{x} \quad \forall \bar{x} y.A(x, y) \rightarrow P(a, f^4 a)}{\Gamma \rightarrow P(a, f^4 a)} \quad \text{cut + c^*}
\]

where \( \psi_1 \) and \( \psi_2 \) are cut-free and \( \psi_2 \) contains only structural and propositional inferences (in \( \psi_2 \) only \( P(f^4 a, a) \) is needed from \( \Gamma \)). The quantifier inferences in \( \psi_1 \) use exactly the substitutions encoded in \( U_1 \) and \( U_2 \). So we have \( |\psi|_q = 10 \).

## 5 Implementation and Experiments

Summing up the previous sections, the structure of our cut-introduction algorithm is the following:

**Algorithm 3 Cut-Introduction**

**Require:** \( \pi \): cut-free proof

\[
T \leftarrow \text{extractTermSet}(\pi)
\]

\[
D \leftarrow \text{getMinimalDecomposition}(T)
\]

\[
C(\bar{x}) \leftarrow \text{getCanonicalSolution}(D)
\]

\[
F(\bar{x}) \leftarrow \text{improveSolution}(C(\bar{x}))
\]

\[
\text{return constructProof}(F(\bar{x}))
\]

Depending on whether the input proof \( \pi \) contains equality reasoning or not we either work modulo quasi-tautologies as described in this paper or modulo tautologies (as described in [8, 7]) in improveSolution and constructProof. In getMinimalDecomposition we can either compute decompositions with a single variable as in [8, 7] or with an unbounded number of variables as described in Section [8]. We denote these two variants with \( \text{CI}^1 \) and \( \text{CI}^* \) respectively.

These algorithms have been implemented in the gapt-system\(^3\) which is a framework for transforming and analyzing formal proofs. It is implemented in Scala and contains data structures such as formulas, sequents, resolution

\(^3\) Generic Architecture for Proof Transformations, [http://www.logic.at/gapt/](http://www.logic.at/gapt/)
and sequent calculus proofs and algorithms like unification, skolemization, cut-elimination as well as backends for several external solvers and provers. For deciding whether a quantifier-free formula is a tautology we use MiniSat\(^4\). We use veriT\(^5\) for deciding whether a quantifier-free formula is a quasi-tautology and prover9\(^6\) for the actual proof construction based on the import described in \[9\].

We have conducted experiments on the prover9-part of the TSTP-library (Thousands of Solutions of Theorem Provers, see \[10\]). The choice of prover9 was motivated by the simple and clean proof output format Ivy which makes proof import (comparatively) easy. This library contains 6341 resolution proofs. Of those, 5254 can be parsed and transformed into a sequent calculus proof using the transformation described in \[9\]. Of those, 2849 have non-trivial termsets (we call a term set trivial if every quantified formula in the end-sequent is instantiated at most once).

The input data we have used for our experiments is this collection of proofs with non-trivial term sets. In this collection 66% use equality reasoning and hence must be treated with the method introduced in this paper. The average term set size is 37,1 but 46% have a term set of size ≤ 10. The experiments have been conducted with version 1.6 of gapt on an Intel i5 QuadCore with 3,33GHz with an allocation of 2GB heap space and a timeout of 60 seconds for the cut-introduction algorithm.

On 19% of the input proofs our algorithm terminates with finding a compression, i.e. a non-trivial decomposition (of size at most that of the original termset) and a proof with cut that realizes this decomposition. On 49% it terminates de-

\(^4\) [http://minisat.se/](http://minisat.se/)
\(^5\) [http://www.verit-solver.org/](http://www.verit-solver.org/)
\(^6\) [http://www.cs.unm.edu/~mccune/prover9/](http://www.cs.unm.edu/~mccune/prover9/)
terminating that the proof is uncompressible, more precisely: that there is no proof with a single ∀-cut which (by cut-elimination) reduces to the given input term set and is of smaller quantifier complexity, see [7]. Figure 2 depicts the return status (in percent) depending on the size of the term set. When reading this figure one should keep in mind the relatively high number of small proofs (see above). One can observe that proofs with term sets up to a size of around 50 can be treated well by our current implementation, beyond that the percentage of timeouts is very large. Small proofs – unsurprisingly – tend to be uncompressible.

In Figure 3 we restrict our attention to runs terminating with a compression. As one can see from the diagram on the left, a significant reduction of quantifier-complexity can be achieved by our method. The diagram on the right demonstrates that forgetful reasoning is highly useful for improving the canonical solution. The points plotted as • are the result after using forgetful resolution only, the points plotted as + are the result after forgetful resolution and paramodulation.

Our experiments also show that the generalization to the introduction of a block of quantifiers introduced in this paper has a strong effect: of the 548 proofs on which CI* finds a compression, 22% are found to be uncompressible by CI¹.

6 Conclusion

We have introduced a cut-introduction method that works modulo equality and is capable of generating cut-formulas containing a block of quantifiers. We have implemented our new method and have conducted a large-scale empirical evaluation which demonstrates its feasibility on realistic examples. Lessons learned from these experiments include that blocks of quantifiers allow for significantly more proofs to be compressed and that forgetful reasoning methods, while rough in theory, are highly useful for our application in practice.
As future work we plan to extend our method to work modulo (suitably specified) equational theories. We also plan to evaluate our method on proofs produced by Tableaux-provers and SMT-solvers. Another important, and non-trivial, extension will be to cope with cuts that contain quantifier-alternations.

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