Decision Procedures in Verification

Propositional Logic (Part 4)

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Until now

1.1 Syntax

1.2 Semantics

1.3 Models, Validity, and Satisfiability

1.4 Normal forms: CNF; DNF; Structure-preserving translation

1.5 Inference Systems and Proofs

1.6 The Propositional Resolution Calculus

1.6 The Propositional Resolution Calculus

Resolution inference rule:

$$\frac{C \vee A \qquad \neg A \vee D}{C \vee D}$$

Terminology: $C \vee D$: resolvent; A: resolved atom

(Positive) factorisation inference rule:

$$\frac{C \vee A \vee A}{C \vee A}$$

The Resolution Calculus *Res*

These are schematic inference rules; for each substitution of the schematic variables C, D, and A, respectively, by propositional clauses and atoms we obtain an inference rule.

As "$\vee$" is considered associative and commutative, we assume that A and $\neg A$ can occur anywhere in their respective clauses.

Soundness of Resolution

Theorem 1.10. Propositional resolution is sound.

Completeness of Resolution

How to show refutational completeness of propositional resolution:

- We have to show: $N \models \bot \;\Rightarrow\; N \vdash_{Res} \bot$,
  or equivalently: If $N \not\vdash_{Res} \bot$, then N has a model.

- Idea: Suppose that we have computed sufficiently many inferences (and not derived $\bot$).

  Now order the clauses in N according to some appropriate ordering, inspect the clauses in ascending order, and construct a series of valuations.

- The limit valuation can be shown to be a model of N.

Stratified Structure of Clause Sets

Let $A \succ B$. Clause sets are then stratified in this form:

$$
\begin{array}{c|c|c|l}
 & & & \\
\hline
 & B \left\{ \begin{array}{c} \quad\dots \lor B \\ \dots \\ \quad\dots \lor B \lor B \\ \dots \\ \neg B \lor \dots \end{array} \right. & & \text{all } D \text{ where } \max(D) = B \\
\hline
\prec \downarrow & \vdots & \vdots & \\
\hline
 & A \left\{ \begin{array}{c} \quad\dots \lor A \\ \dots \\ \quad\dots \lor A \lor A \\ \dots \\ \neg A \lor \dots \\ \dots \end{array} \right. & & \text{all } C \text{ where } \max(C) = A \\
\hline
 & & & \\
\end{array}
$$

Closure of Clause Sets under *Res*

$$
\begin{aligned}
Res(N) &= \{C \mid C \text{ is concl. of a rule in } Res \text{ w/ premises in } N\} \\
Res^0(N) &= N \\
Res^{n+1}(N) &= Res(Res^n(N)) \cup Res^n(N), \text{ for } n \geq 0 \\
Res^*(N) &= \textstyle\bigcup_{n \geq 0} Res^n(N)
\end{aligned}
$$

N is called saturated (wrt. resolution), if $Res(N) \subseteq N$.

Proposition 1.12

(i) $Res^*(N)$ is saturated.

(ii) Res is refutationally complete, iff for each set N of ground clauses:

$$N \models \bot \;\Leftrightarrow\; \bot \in Res^*(N)$$

Construction of Interpretations

Given: set N of clauses, atom ordering $\succ$.

Wanted: Valuation $\mathcal{A}$ such that

- "many" clauses from N are valid in $\mathcal{A}$;
- $\mathcal{A} \models N$, *if* N *is saturated and* $\bot \notin N$.

Construction according to $\succ$, starting with the minimal clause.

Main Ideas of the Construction

- Clauses are considered in the order given by $\prec$. We construct a model for N incrementally.

- When considering C, one already has a partial interpretation I_C (initially $I_C = \emptyset$) available.

  In what follows, instead of referring to partial valuations $\mathcal{A}_C$ we will refer to partial interpretations I_C (the set of atoms which are true in the valuation $\mathcal{A}_C$).

- If C is true in the partial interpretation I_C, nothing is done. $(\Delta_C = \emptyset)$.

- If C is false, one would like to change I_C such that C becomes true.

Main Ideas of the Construction

- Changes should, however, be *monotone*. One never deletes anything from I_C and the truth value of clauses smaller than C should be maintained the way it was in I_C.

- Hence, one chooses $\Delta_C = \{A\}$ if, and only if, C is false in I_C, if A occurs positively in C (*adding A will make C become true*) and if this occurrence in C is strictly maximal in the ordering on literals (*changing the truth value of A has no effect on smaller clauses*).

Factorization Reduces Counterexamples

$$\frac{\neg P_1 \vee \neg P_1 \vee P_3 \vee P_3 \vee P_0}{\neg P_1 \vee \neg P_1 \vee P_3 \vee P_0}$$

Construction of I for the extended clause set:

	clauses C	I_C	Δ_C	Remarks
1	$\neg P_0$	$\emptyset$	$\emptyset$	
2	$P_0 \vee P_1$	$\emptyset$	$\{P_1\}$	
3	$P_1 \vee P_2$	$\{P_1\}$	$\emptyset$	
4	$\neg P_1 \vee P_2$	$\{P_1\}$	$\{P_2\}$	
9	$\neg P_1 \vee \neg P_1 \vee P_3 \vee P_0$	$\{P_1, P_2\}$	$\{P_3\}$	
8	$\neg P_1 \vee \neg P_1 \vee P_3 \vee P_3 \vee P_0$	$\{P_1, P_2, P_3\}$	$\emptyset$	true in $\mathcal{A}_C$
5	$\neg P_1 \vee P_4 \vee P_3 \vee P_0$	$\{P_1, P_2, P_3\}$	$\emptyset$	
6	$\neg P_1 \vee \neg P_4 \vee P_3$	$\{P_1, P_2, P_3\}$	$\emptyset$	true in $\mathcal{A}_C$
7	$\neg P_3 \vee P_5$	$\{P_1, P_2, P_3\}$	$\{P_5\}$	

The resulting $I = \{P_1, P_2, P_3, P_5\}$ is a model of the clause set.

Construction of Candidate Models Formally

Let $N, \succ$ be given. We define sets I_C and Δ_C for all ground clauses C over the given signature inductively over $\succ$:

$$I_C := \bigcup_{C \succ D} \Delta_D$$

$$\Delta_C := \begin{cases} \{A\}, & \text{if } C \in N,\ C = C' \vee A,\ A \succ C',\ I_C \not\models C \\ \emptyset, & \text{otherwise} \end{cases}$$

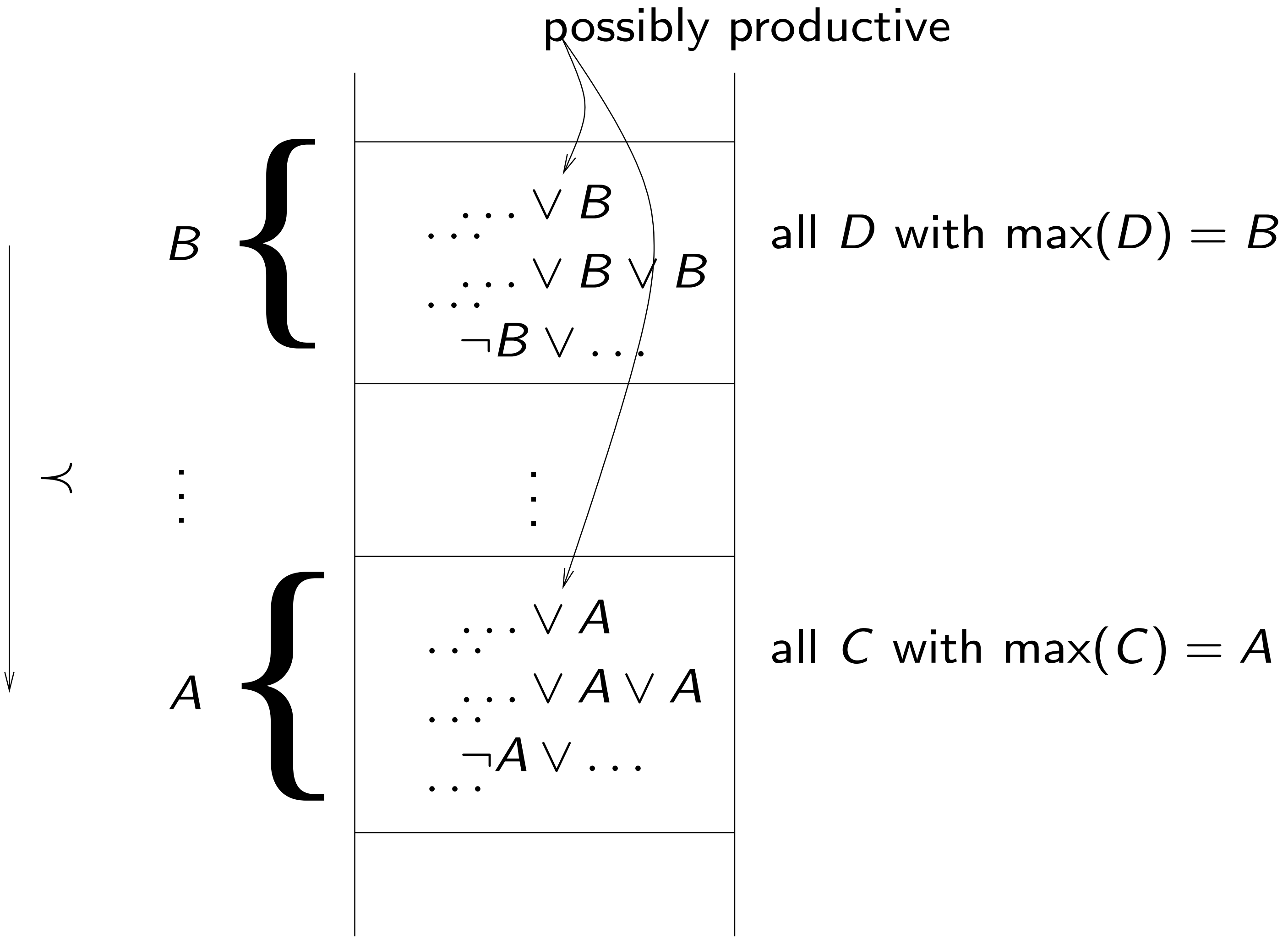
We say that C produces A, if $\Delta_C = \{A\}$.

The candidate model for N (wrt. $\succ$) is given as $I_N^{\succ} := \bigcup_C \Delta_C$.

We also simply write I_N, or I, for $I_N^{\succ}$ if $\succ$ is either irrelevant or known from the context.

Structure of $N, \succ$

Let $A \succ B$; producing a new atom does not affect smaller clauses.



Some Properties of the Construction

Proposition 1.13:

(i) $C = \neg A \vee C' \ \Rightarrow\ $ no $D \succeq C$ produces A.

(ii) C productive $\Rightarrow I_C \cup \Delta_C \models C$.

(iii) Let $D' \succ D \succeq C$. Then

$$I_D \cup \Delta_D \models C \Rightarrow I_{D'} \cup \Delta_{D'} \models C \text{ and } I_N \models C.$$

If, in addition, $C \in N$ or $\max(D) \succ \max(C)$:

$$I_D \cup \Delta_D \not\models C \Rightarrow I_{D'} \cup \Delta_{D'} \not\models C \text{ and } I_N \not\models C.$$

Some Properties of the Construction

(iv) Let $D' \succ D \succ C$. Then

$$I_D \models C \Rightarrow I_{D'} \models C \text{ and } I_N \models C.$$

If, in addition, $C \in N$ or $\max(D) \succ \max(C)$:

$$I_D \not\models C \Rightarrow I_{D'} \not\models C \text{ and } I_N \not\models C.$$

(v) $D = C \vee A$ produces $A \Rightarrow I_N \not\models C$.

Model Existence Theorem

Theorem 1.14 (Bachmair & Ganzinger):
Let $\succ$ be a clause ordering, let N be saturated wrt. *Res*, and suppose that $\bot \notin N$. Then $I_N^{\succ} \models N$.

Corollary 1.15:
Let N be saturated wrt. *Res*. Then $N \models \bot \Leftrightarrow \bot \in N$.

Model Existence Theorem

Proof:

Suppose $\bot \notin N$, but $I_N^{\succ} \not\models N$. Let $C \in N$ minimal (in $\succ$) such that $I_N^{\succ} \not\models C$. Since C is false in I_N, C is not productive. As $C \neq \bot$ there exists a maximal atom A in C.

Case 1: $C = \neg A \vee C'$ (i.e., the maximal atom occurs negatively)
$\Rightarrow I_N \models A$ and $I_N \not\models C'$
$\Rightarrow$ some $D = D' \vee A \in N$ produces A. As $\frac{D' \vee A \qquad\qquad \neg A \vee C'}{D' \vee C'}$, we infer that $D' \vee C' \in N$, and $C \succ D' \vee C'$ and $I_N \not\models D' \vee C'$
$\Rightarrow$ contradicts minimality of C.

Case 2: $C = C' \vee A \vee A$. Then $\frac{C' \vee A \vee A}{C' \vee A}$ yields a smaller counterexample $C' \vee A \in N$. $\Rightarrow$ contradicts minimality of C.

Ordered Resolution with Selection

Ideas for improvement:

1. In the completeness proof (Model Existence Theorem) one only needs to resolve and factor maximal atoms
   $\Rightarrow$ if the calculus is restricted to inferences involving maximal atoms, the proof remains correct
   $\Rightarrow$ *order restrictions*

2. In the proof, it does not really matter with which negative literal an inference is performed
   $\Rightarrow$ choose a negative literal don't-care-nondeterministically
   $\Rightarrow$ *selection*

Selection Functions

A selection function is a mapping

$$S : C \quad \mapsto \quad \text{set of occurrences of } \textit{negative} \text{ literals in } C$$

Example of selection with selected literals indicated as $\boxed{X}$:

$$\boxed{\neg A} \vee \neg A \vee B$$

$$\boxed{\neg B_0} \vee \boxed{\neg B_1} \vee A$$

Ordered resolution

In the completeness proof, we talk about (strictly) maximal literals of clauses.

Resolution Calculus $Res_S^{\succ}$

$$\frac{C \vee A \qquad D \vee \neg A}{C \vee D}$$ [ordered resolution with selection]

if

(i) $A \succ C$;

(ii) nothing is selected in C by S;

(iii) $\neg A$ is selected in $D \vee \neg A$,
or else nothing is selected in $D \vee \neg A$ and $\neg A \succeq \max(D)$.

Note: For positive literals, $A \succ C$ is the same as $A \succ \max(C)$.

Resolution Calculus $Res_S^{\succ}$

$$\frac{C \lor A \lor A}{(C \lor A)} \qquad \text{[ordered factoring]}$$

if A is maximal in C and nothing is selected in C.

Search Spaces Become Smaller

1	$A \vee B$	
2	$A \vee \boxed{\neg B}$	
3	$\neg A \vee B$	
4	$\neg A \vee \boxed{\neg B}$	
5	$B \vee B$	Res 1, 3
6	B	Fact 5
7	$\neg A$	Res 6, 4
8	A	Res 6, 2
9	$\perp$	Res 8, 7

we assume $A \succ B$ and S as indicated by $\boxed{X}$. The maximal literal in a clause is depicted in red.

With this ordering and selection function the refutation proceeds strictly deterministically in this example. Generally, proof search will still be non-deterministic but the search space will be much smaller than with unrestricted resolution.

1.7 The DPLL Procedure

Goal:

Given a propositional formula in CNF (or alternatively, a finite set N of clauses), check whether it is satisfiable (and optionally: output *one* solution, if it is satisfiable).

Satisfiability of Clause Sets

$\mathcal{A} \models N$ if and only if $\mathcal{A} \models C$ for all clauses C in N.

$\mathcal{A} \models C$ if and only if $\mathcal{A} \models L$ for some literal $L \in C$.

Partial Valuations

Since we will construct satisfying valuations incrementally, we consider partial valuations (that is, partial mappings $\mathcal{A} : \Pi \rightarrow \{0, 1\}$).

We start with an empty valuation and try to extend it step by step to all variables occurring in N.

If $\mathcal{A}$ is a partial valuation, then literals and clauses can be true, false, or undefined under $\mathcal{A}$.

A clause is true under $\mathcal{A}$ if one of its literals is true; it is false (or "conflicting") if all its literals are false; otherwise it is undefined (or "unresolved").

Unit Clauses

Observation:

Let $\mathcal{A}$ be a partial valuation. If the set N contains a clause C, such that all literals but one in C are false under $\mathcal{A}$, then the following properties are equivalent:

- there is a valuation that is a model of N and extends $\mathcal{A}$.
- there is a valuation that is a model of N and extends $\mathcal{A}$ and makes the remaining literal L of C true.

C is called a unit clause; L is called a unit literal.

Pure Literals

One more observation:

Let $\mathcal{A}$ be a partial valuation and P a variable that is undefined under $\mathcal{A}$. If P occurs only positively (or only negatively) in the unresolved clauses in N, then the following properties are equivalent:

- there is a valuation that is a model of N and extends $\mathcal{A}$.
- there is a valuation that is a model of N and extends $\mathcal{A}$ and assigns true (false) to P.

P is called a pure literal.

The Davis-Putnam-Logemann-Loveland Proc.

```
boolean DPLL(clause set N, partial valuation 𝒜) {
    if (all clauses in N are true under 𝒜) return true;
    elsif (some clause in N is false under 𝒜) return false;
    elsif (N contains unit clause P) return DPLL(N, 𝒜 ∪ {P ↦ 1});
    elsif (N contains unit clause ¬P) return DPLL(N, 𝒜 ∪ {P ↦ 0});
    elsif (N contains pure literal P) return DPLL(N, 𝒜 ∪ {P ↦ 1});
    elsif (N contains pure literal ¬P) return DPLL(N, 𝒜 ∪ {P ↦ 0});
    else {
        let P be some undefined variable in N;
        if (DPLL(N, 𝒜 ∪ {P ↦ 0})) return true;
        else return DPLL(N, 𝒜 ∪ {P ↦ 1});
    }
}
```

The Davis-Putnam-Logemann-Loveland Proc.

Initially, DPLL is called with the clause set N and with an empty partial valuation $\mathcal{A}$.

The Davis-Putnam-Logemann-Loveland Proc.

In practice, there are several changes to the procedure:

The pure literal check is often omitted (it is too expensive).

The branching variable is not chosen randomly.

The algorithm is implemented iteratively;
the backtrack stack is managed explicitly
(it may be possible and useful to backtrack more than one level).

DPLL Iteratively

An iterative (and generalized) version:

```
status = preprocess();
if (status != UNKNOWN) return status;
while(1) {
    decide_next_branch();
    while(1) {
        status = deduce();
        if (status == CONFLICT) {
            blevel = analyze_conflict();
            if (blevel == 0) return UNSATISFIABLE;
            else backtrack(blevel); }
        else if (status == SATISFIABLE) return SATISFIABLE;
        else break;
    }
}
```

DPLL Iteratively

`preprocess()`

preprocess the input (as far as it is possible without branching);
return CONFLICT or SATISFIABLE or UNKNOWN.

`decide_next_branch()`

choose the right undefined variable to branch;
decide whether to set it to 0 or 1;
increase the backtrack level.

DPLL Iteratively

`deduce()`

make further assignments to variables (e.g., using the unit clause rule) until a satisfying assignment is found, or until a conflict is found, or until branching becomes necessary; return CONFLICT or SATISFIABLE or UNKNOWN.

DPLL Iteratively

`analyze_conflict()`

check where to backtrack.

`backtrack(blevel)`

backtrack to `blevel`;

flip the branching variable on that level;

undo the variable assignments in between.

Branching Heuristics

Choosing the right undefined variable to branch is important for efficiency, but the branching heuristics may be expensive itself.

State of the art: use branching heuristics that need not be recomputed too frequently.

In general: choose variables that occur frequently.

The Deduction Algorithm

For applying the unit rule, we need to know the number of literals in a clause that are not false.

Maintaining this number is expensive, however.

The Deduction Algorithm

Better approach: "Two watched literals":

In each clause, select two (currently undefined) "watched" literals.

For each variable P, keep a list of all clauses in which P is watched and a list of all clauses in which $\neg P$ is watched.

If an undefined variable is set to 0 (or to 1), check all clauses in which P (or $\neg P$) is watched and watch another literal (that is true or undefined) in this clause if possible.

Watched literal information need not be restored upon backtracking.

Conflict Analysis and Learning

Goal: Reuse information that is obtained in one branch in further branches.

Method: Learning:

> If a conflicting clause is found, use the resolution rule to derive a new clause and add it to the current set of clauses.

Problem: This may produce a large number of new clauses; therefore it may become necessary to delete some of them afterwards to save space.

Backjumping

Related technique:

non-chronological backtracking ("backjumping"):

If a conflict is independent of some earlier branch, try to skip that over that backtrack level.

Restart

Runtimes of DPLL-style procedures depend extremely on the choice of branching variables.

If no solution is found within a certain time limit, it can be useful to restart from scratch with another choice of branchings (but learned clauses may be kept).

A succinct formulation

State: $M||F$,

where:

- M partial assignment (sequence of literals),

  some literals are annotated (L^d: decision literal)

- F clause set.

A succinct formulation

UnitPropagation

$M||F, C \vee L \Rightarrow M, L||F, C \vee L$ if $M \models \neg C$, and L undef. in M

Decide

$M||F \Rightarrow M, L^d||F$ if L or $\neg L$ occurs in F, L undef. in M

Fail

$M||F, C \Rightarrow \text{Fail}$ if $M \models \neg C$, M contains no decision literals

Backjump

$$M, L^d, N||F \Rightarrow M, L'||F \quad \text{if} \begin{cases} \text{there is some clause } C \vee L' \text{ s.t.:} \\ F \models C \vee L', M \models \neg C, \\ L' \text{ undefined in } M \\ L' \text{ or } \neg L' \text{ occurs in } F. \end{cases}$$

Example

Assignment:	Clause set:	
$\emptyset$	$\|\|\neg P_1 \vee P_2, \neg P_3 \vee P_4, \neg P_5 \vee \neg P_6, P_6 \vee \neg P_5 \vee \neg P_2$	$\Rightarrow$ (Decide)
P_1	$\|\|\neg P_1 \vee P_2, \neg P_3 \vee P_4, \neg P_5 \vee \neg P_6, P_6 \vee \neg P_5 \vee \neg P_2$	$\Rightarrow$ (UnitProp)
$P_1 P_2$	$\|\|\neg P_1 \vee P_2, \neg P_3 \vee P_4, \neg P_5 \vee \neg P_6, P_6 \vee \neg P_5 \vee \neg P_2$	$\Rightarrow$ (Decide)
$P_1 P_2 P_3$	$\|\|\neg P_1 \vee P_2, \neg P_3 \vee P_4, \neg P_5 \vee \neg P_6, P_6 \vee \neg P_5 \vee \neg P_2$	$\Rightarrow$ (UnitProp)
$P_1 P_2 P_3 P_4$	$\|\|\neg P_1 \vee P_2, \neg P_3 \vee P_4, \neg P_5 \vee \neg P_6, P_6 \vee \neg P_5 \vee \neg P_2$	$\Rightarrow$ (Decide)
$P_1 P_2 P_3 P_4 P_5$	$\|\|\neg P_1 \vee P_2, \neg P_3 \vee P_4, \neg P_5 \vee \neg P_6, P_6 \vee \neg P_5 \vee \neg P_2$	$\Rightarrow$ (UnitProp)
$P_1 P_2 P_3 P_4 P_5 \neg P_6$	$\|\|\neg P_1 \vee P_2, \neg P_3 \vee P_4, \neg P_5 \vee \neg P_6, P_6 \vee \neg P_5 \vee \neg P_2$	$\Rightarrow$ (Backtrack)
$P_1 P_2 P_3 P_4 \neg P_5$	$\|\|\neg P_1 \vee P_2, \neg P_3 \vee P_4, \neg P_5 \vee \neg P_6, P_6 \vee \neg P_5 \vee \neg P_2$	...

DPLL with learning

The DPLL system with learning consists of the four transition rules of the Basic DPLL system, plus the following two additional rules:

Learn

$M||F \Rightarrow M||F, C$ if all atoms of C occur in F and $F \models C$

Forget

$M||F, C \Rightarrow M||F$ if $F \models C$

In these two rules, the clause C is said to be learned and forgotten, respectively.

Further Information

The ideas described so far heve been implemented in the SAT checker Chaff.

Further information:
Lintao Zhang and Sharad Malik:
The Quest for Efficient Boolean Satisfiability Solvers,
Proc. CADE-18, LNAI 2392, pp. 295–312, Springer, 2002.