Decision Procedures for Verification

Part 1. Propositional Logic (2)

4.11.2014

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Last time

1.1 Syntax

- Language
  - propositional variables
  - logical symbols

    $\Rightarrow$ Boolean combinations
- Propositional Formulae

1.2 Semantics

- Valuations
- Truth value of a formula in a valuation
- Models, Validity, and Satisfiability

1.3 Models, Validity, and Satisfiability

F is valid in $\mathcal{A}$ ($\mathcal{A}$ is a model of F; F holds under $\mathcal{A}$):

$$\mathcal{A} \models F :\Leftrightarrow \mathcal{A}(F) = 1$$

F is valid (or is a tautology):

$$\models F :\Leftrightarrow \mathcal{A} \models F \text{ for all } \Pi\text{-valuations } \mathcal{A}$$

F is called satisfiable iff there exists an $\mathcal{A}$ such that $\mathcal{A} \models F$.

Otherwise F is called unsatisfiable (or contradictory).

Entailment and Equivalence

F entails (implies) G (or G is a consequence of F), written $F \models G$, if for all Π-valuations $\mathcal{A}$, whenever $\mathcal{A} \models F$ then $\mathcal{A} \models G$.

F and G are called equivalent if for all Π-valuations $\mathcal{A}$ we have $\mathcal{A} \models F \Leftrightarrow \mathcal{A} \models G$.

Proposition 1.1:
F entails G iff $(F \rightarrow G)$ is valid

Proposition 1.2:
F and G are equivalent iff $(F \leftrightarrow G)$ is valid.

Entailment and Equivalence

Extension to sets of formulas N in the "natural way", e.g., $N \models F$ if for all Π-valuations $\mathcal{A}$: if $\mathcal{A} \models G$ for all $G \in N$, then $\mathcal{A} \models F$.

Validity vs. Unsatisfiability

Validity and unsatisfiability are just two sides of the same medal as explained by the following proposition.

Proposition 1.3:

$$F \text{ valid} \Leftrightarrow \neg F \text{ unsatisfiable}$$

Hence in order to design a theorem prover (validity checker) it is sufficient to design a checker for unsatisfiability.

Q: In a similar way, entailment $N \models F$ can be reduced to unsatisfiability. How?

Validity vs. Unsatisfiability

Validity and unsatisfiability are just two sides of the same medal as explained by the following proposition.

Proposition 1.4:

$$N \models F \Leftrightarrow N \cup \neg F \text{ unsatisfiable}$$

Hence in order to design a theorem prover (validity/entailment checker) it is sufficient to design a checker for unsatisfiability.

Checking Unsatisfiability

Every formula F contains only finitely many propositional variables. Obviously, $\mathcal{A}(F)$ depends only on the values of those finitely many variables in F under $\mathcal{A}$.

If F contains n distinct propositional variables, then it is sufficient to check 2^n valuations to see whether F is satisfiable or not.
$\Rightarrow$ truth table.

So the satisfiability problem is clearly decidable
(but, by Cook's Theorem, NP-complete).

Nevertheless, in practice, there are (much) better methods than truth tables to check the satisfiability of a formula. (later more)

Checking Unsatisfiability

The satisfiability problem is clearly decidable
(but, by Cook's Theorem, NP-complete).

For sets of propositional formulae of a certain type, satisfiability can be checked in polynomial time:

Examples: 2SAT, Horn-SAT (will be discussed in the exercises)

Dichotomy theorem. Schaefer [Schaefer, STOC 1978] identified six classes of sets S of Boolean formulae for which $SAT(S)$ is in PTIME. He proved that all other types of sets of formulae yield an NP-complete problem.

Substitution Theorem

Proposition 1.5:

Let F and G be equivalent formulas, let H be a formula in which F occurs as a subformula.

Then H is equivalent to H' where H' is obtained from H by replacing the occurrence of the subformula F by G.
(Notation: $H = H[F]$, $H' = H[G]$.)

Proof: By induction over the formula structure of H.

Structural Induction

Goal: Prove a property P of propositional formulae

Prove that for every formula F, $P(F)$ holds.

Induction basis: Show that $P(F)$ holds for all $F \in \Pi \cup \{\top, \bot\}$

Let F be a formula (not in $\Pi \cup \{\top, \bot\}$).

Induction hypothesis: We assume that $P(G)$ holds for all strict subformulae G of F.

Induction step: Using the induction hypothesis, we show that $P(F)$ holds as well.
In order to prove that $P(F)$ holds we usually need to consider various cases (reflecting the way the formula F is built):

Case 1: $F = \neg G$
Case 2: $F = G_1 \wedge G_2$
Case 3: $F = G_1 \vee G_2$
Case 4: $F = G_1 \rightarrow G_2$
Case 5: $F = G_1 \leftrightarrow G_2$

Some Important Equivalences

Proposition 1.6:

The following equivalences are valid for all formulas F, G, H:

$$(F \wedge F) \leftrightarrow F$$

$$(F \vee F) \leftrightarrow F \qquad \text{(Idempotency)}$$

$$(F \wedge G) \leftrightarrow (G \wedge F)$$

$$(F \vee G) \leftrightarrow (G \vee F) \qquad \text{(Commutativity)}$$

$$(F \wedge (G \wedge H)) \leftrightarrow ((F \wedge G) \wedge H)$$

$$(F \vee (G \vee H)) \leftrightarrow ((F \vee G) \vee H) \qquad \text{(Associativity)}$$

$$(F \wedge (G \vee H)) \leftrightarrow ((F \wedge G) \vee (F \wedge H))$$

$$(F \vee (G \wedge H)) \leftrightarrow ((F \vee G) \wedge (F \vee H)) \qquad \text{(Distributivity)}$$

Some Important Equivalences

Proposition 1.7:

The following equivalences are valid for all formulas F, G, H:

$$(F \wedge (F \vee G)) \leftrightarrow F$$

$$(F \vee (F \wedge G)) \leftrightarrow F \qquad \text{(Absorption)}$$

$$(\neg\neg F) \leftrightarrow F \qquad \text{(Double Negation)}$$

$$\neg(F \wedge G) \leftrightarrow (\neg F \vee \neg G)$$

$$\neg(F \vee G) \leftrightarrow (\neg F \wedge \neg G) \qquad \text{(De Morgan's Laws)}$$

$$(F \wedge G) \leftrightarrow F, \text{ if } G \text{ is a tautology}$$

$$(F \vee G) \leftrightarrow \top, \text{ if } G \text{ is a tautology} \qquad \text{(Tautology Laws)}$$

$$(F \wedge G) \leftrightarrow \bot, \text{ if } G \text{ is unsatisfiable}$$

$$(F \vee G) \leftrightarrow F, \text{ if } G \text{ is unsatisfiable} \qquad \text{(Tautology Laws)}$$

1.4 Normal Forms

We define conjunctions of formulas as follows:

$\bigwedge_{i=1}^{0} F_i = \top$.

$\bigwedge_{i=1}^{1} F_i = F_1$.

$\bigwedge_{i=1}^{n+1} F_i = \bigwedge_{i=1}^{n} F_i \wedge F_{n+1}$.

and analogously disjunctions:

$\bigvee_{i=1}^{0} F_i = \bot$.

$\bigvee_{i=1}^{1} F_i = F_1$.

$\bigvee_{i=1}^{n+1} F_i = \bigvee_{i=1}^{n} F_i \vee F_{n+1}$.

Literals and Clauses

A literal is either a propositional variable P or a negated propositional variable $\neg P$.

A clause is a (possibly empty) disjunction of literals.

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A clause is a (possibly empty) disjunction of literals.

Example of clauses:

$\perp$	the empty clause
P	positive unit clause
$\neg P$	negative unit clause
$P \vee Q \vee R$	positive clause
$P \vee \neg Q \vee \neg R$	clause
$P \vee P \vee \neg Q \vee \neg R \vee R$	allow repetitions/complementary literals

CNF and DNF

A formula is in conjunctive normal form (CNF, clause normal form), if it is a conjunction of disjunctions of literals (or in other words, a conjunction of clauses).

A formula is in disjunctive normal form (DNF), if it is a disjunction of conjunctions of literals.

Warning: definitions in the literature differ:

are complementary literals permitted?
are duplicated literals permitted?
are empty disjunctions/conjunctions permitted?

CNF and DNF

Checking the validity of CNF formulas or the unsatisfiability of DNF formulas is easy:

A formula in CNF is valid, if and only if each of its disjunctions contains a pair of complementary literals P and $\neg P$.

Conversely, a formula in DNF is unsatisfiable, if and only if each of its conjunctions contains a pair of complementary literals P and $\neg P$.

On the other hand, checking the unsatisfiability of CNF formulas or the validity of DNF formulas is known to be coNP-complete.

Conversion to CNF/DNF

Proposition 1.8:

For every formula there is an equivalent formula in CNF (and also an equivalent formula in DNF).

Proof:

We consider the case of CNF.

Apply the following rules as long as possible (modulo associativity and commutativity of $\wedge$ and $\vee$):

Step 1: Eliminate equivalences:

$$(F \leftrightarrow G) \;\Rightarrow_K\; (F \rightarrow G) \wedge (G \rightarrow F)$$

Conversion to CNF/DNF

Step 2: Eliminate implications:

$$(F \rightarrow G) \;\Rightarrow_K\; (\neg F \vee G)$$

Step 3: Push negations downward:

$$\neg(F \vee G) \;\Rightarrow_K\; (\neg F \wedge \neg G)$$
$$\neg(F \wedge G) \;\Rightarrow_K\; (\neg F \vee \neg G)$$

Step 4: Eliminate multiple negations:

$$\neg\neg F \;\Rightarrow_K\; F$$

The formula obtained from a formula F after applying steps 1-4 is called the negation normal form (NNF) of F

Conversion to CNF/DNF

Step 5: Push disjunctions downward:

$$(F \wedge G) \vee H \;\Rightarrow_K\; (F \vee H) \wedge (G \vee H)$$

Step 6: Eliminate $\top$ and $\bot$:

$$\begin{aligned}
(F \wedge \top) &\;\Rightarrow_K\; F \\
(F \wedge \bot) &\;\Rightarrow_K\; \bot \\
(F \vee \top) &\;\Rightarrow_K\; \top \\
(F \vee \bot) &\;\Rightarrow_K\; F \\
\neg\bot &\;\Rightarrow_K\; \top \\
\neg\top &\;\Rightarrow_K\; \bot
\end{aligned}$$

Conversion to CNF/DNF

Proving termination is easy for most of the steps; only steps 1, 3 and 5 are a bit more complicated.

The resulting formula is equivalent to the original one and in CNF.

The conversion of a formula to DNF works in the same way, except that disjunctions have to be pushed downward in step 5.

Complexity

Conversion to CNF (or DNF) may produce a formula whose size is exponential in the size of the original one.

Satisfiability-preserving Transformations

The goal

“find a formula G in CNF such that $\models F \leftrightarrow G$”

is unpractical.

But if we relax the requirement to

“find a formula G in CNF such that $F \models \bot$ iff $G \models \bot$”

we can get an efficient transformation.

Satisfiability-preserving Transformations

Idea:
A formula $F[F']$ is satisfiable iff $F[P] \wedge (P \leftrightarrow F')$ is satisfiable (where P new propositional variable that works as abbreviation for F').

We can use this rule recursively for all subformulas in the original formula (this introduces a linear number of new propositional variables).

Conversion of the resulting formula to CNF increases the size only by an additional factor (each formula $P \leftrightarrow F'$ gives rise to at most one application of the distributivity law).

Optimized Transformations

A further improvement is possible by taking the polarity of the subformula F into account.

Assume that F contains neither $\rightarrow$ nor $\leftrightarrow$. A subformula F' of F has positive polarity in F, if it occurs below an even number of negation signs; it has negative polarity in F, if it occurs below an odd number of negation signs.

Optimized Transformations

Proposition 1.9:
Let $F[F']$ be a formula containing neither $\rightarrow$ nor $\leftrightarrow$; let P be a propositional variable not occurring in $F[F']$.

If F' has positive polarity in F, then $F[F']$ is satisfiable if and only if $F[P] \wedge (P \rightarrow F')$ is satisfiable.

If F' has negative polarity in F, then $F[F']$ is satisfiable if and only if $F[P] \wedge (F' \rightarrow P)$ is satisfiable.

Proof:
Exercise.

This satisfiability-preserving transformation to clause form is also called structure-preserving transformation to clause form.

Optimized Transformations

Example: Let $F = (Q_1 \wedge Q_2) \vee (R_1 \wedge R_2)$.

The following are equivalent:

- $F \models \perp$
- $P_F \;\wedge\; (P_F \leftrightarrow (P_{Q_1 \wedge Q_2} \vee P_{R_1 \wedge R_2}) \;\wedge\; (P_{Q_1 \wedge Q_2} \leftrightarrow (Q_1 \wedge Q_2))$
  $\wedge\; (P_{R_1 \wedge R_2} \leftrightarrow (R_1 \wedge R_2)) \models \perp$
- $P_F \;\wedge\; (P_F \rightarrow (P_{Q_1 \wedge Q_2} \vee P_{R_1 \wedge R_2}) \;\wedge\; (P_{Q_1 \wedge Q_2} \rightarrow (Q_1 \wedge Q_2))$
  $\wedge\; (P_{R_1 \wedge R_2} \rightarrow (R_1 \wedge R_2)) \models \perp$
- $P_F \;\wedge\; (\neg P_F \vee P_{Q_1 \wedge Q_2} \vee P_{R_1 \wedge R_2}) \;\wedge\; (\neg P_{Q_1 \wedge Q_2} \vee Q_1) \wedge (\neg P_{Q_1 \wedge Q_2} \vee Q_2)$
  $\wedge\; (\neg P_{R_1 \wedge R_2} \vee R_1) \wedge (\neg P_{R_1 \wedge R_2} \vee R_2)) \models$

Decision Procedures for Satisfiability

- Simple Decision Procedures

  truth table method

- The Resolution Procedure

- The Davis-Putnam-Logemann-Loveland Algorithm

1.5 Inference Systems and Proofs

Inference systems Γ (proof calculi) are sets of tuples

$$(F_1, \ldots, F_n, F_{n+1}), \;\; n \geq 0,$$

called inferences or inference rules, and written

$$\frac{\overbrace{F_1 \;\; \ldots \;\; F_n}^{\text{premises}}}{\underbrace{F_{n+1}}_{\text{conclusion}}} \;\; .$$

Clausal inference system: premises and conclusions are clauses. One also considers inference systems over other data structures.

Proofs

A proof in Γ of a formula F from a a set of formulas N (called assumptions) is a sequence $F_1, \dots, F_k$ of formulas where

(i) $F_k = F$,

(ii) for all $1 \le i \le k$: $F_i \in N$, or else there exists an inference $(F_{i_1}, \dots, F_{i_{n_i}}, F_i)$ in Γ, such that $0 \le i_j < i$, for $1 \le j \le n_i$.

Soundness and Completeness

Provability $\vdash_\Gamma$ of F from N in Γ:

$N \vdash_\Gamma F \quad :\Leftrightarrow \quad$ there exists a proof Γ of F from N.

Γ is called sound $\quad :\Leftrightarrow$

$$\frac{F_1 \ \ \ldots \ \ F_n}{F} \in \Gamma \quad \Rightarrow \quad F_1, \ldots, F_n \models F$$

Γ is called complete $\quad :\Leftrightarrow$

$$N \models F \quad \Rightarrow \quad N \vdash_\Gamma F$$

Γ is called refutationally complete $\quad :\Leftrightarrow$

$$N \models \bot \quad \Rightarrow \quad N \vdash_\Gamma \bot$$

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.

Sample Refutation

1.	$\neg P \vee \neg P \vee Q$	(given)
2.	$P \vee Q$	(given)
3.	$\neg R \vee \neg Q$	(given)
4.	R	(given)
5.	$\neg P \vee Q \vee Q$	(Res. 2. into 1.)
6.	$\neg P \vee Q$	(Fact. 5.)
7.	$Q \vee Q$	(Res. 2. into 6.)
8.	Q	(Fact. 7.)
9.	$\neg R$	(Res. 8. into 3.)
10.	$\perp$	(Res. 4. into 9.)

Resolution with Implicit Factorization *RIF*

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

1.	$\neg P \vee \neg P \vee Q$	(given)
2.	$P \vee Q$	(given)
3.	$\neg R \vee \neg Q$	(given)
4.	R	(given)
5.	$\neg P \vee Q \vee Q$	(Res. 2. into 1.)
6.	$Q \vee Q \vee Q$	(Res. 2. into 5.)
7.	$\neg R$	(Res. 6. into 3.)
8.	$\perp$	(Res. 4. into 7.)

Soundness of Resolution

Theorem 1.10. Propositional resolution is sound.

Proof:

Let $\mathcal{A}$ valuation. To be shown:

(i) for resolution: $\mathcal{A} \models C \vee A,\ \mathcal{A} \models D \vee \neg A \ \Rightarrow\ \mathcal{A} \models C \vee D$

(ii) for factorization: $\mathcal{A} \models C \vee A \vee A \ \Rightarrow\ \mathcal{A} \models C \vee A$

(i): Assume $\mathcal{A}^*(C \vee A) = 1, \mathcal{A}^*(D \vee \neg A) = 1$.
Two cases need to be considered: (a) $\mathcal{A}^*(A) = 1$, or (b) $\mathcal{A}^*(\neg A) = 1$.

(a) $\mathcal{A} \models A \Rightarrow \mathcal{A} \models D \Rightarrow \mathcal{A} \models C \vee D$

(b) $\mathcal{A} \models \neg A \Rightarrow \mathcal{A} \models C \Rightarrow \mathcal{A} \models C \vee D$

(ii): Assume $\mathcal{A} \models C \vee A \vee A$. Note that $\mathcal{A}^*(C \vee A \vee A) = \mathcal{A}^*(C \vee A)$, i.e. the conclusion is also true in $\mathcal{A}$.

Soundness of Resolution

Note: In propositional logic we have:

1. $\mathcal{A} \models L_1 \vee \ldots \vee L_n \;\Leftrightarrow\;$ there exists i: $\mathcal{A} \models L_i$.
2. $\mathcal{A} \models A$ or $\mathcal{A} \models \neg A$.

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.