

Lecture 7

The compactness theorem

Print version of the lecture in *Introduction to Logic for Computer Science*

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These lecture notes are very minor variants of the ones made by Prof James Worrell and Dr Christoph Haase for their 'Logic and Proof' course at Oxford.

7.1

1 The compactness theorem

In this lecture we prove a fundamental result about propositional logic called the *Compactness Theorem*. This will play an important role in the second half of the course when we study predicate logic. This is due to our use of *Herbrand's Theorem* to reduce reasoning about formulas of predicate logic to reasoning about infinite sets of formulas of propositional logic.

Before stating and proving the Compactness Theorem we need to introduce one new piece of terminology. A *partial assignment* is a function $\mathcal{A} : D \rightarrow \{0, 1\}$, where $D \subseteq \{p_1, p_2, \dots\}$ is a set of propositional variables.¹ The set D is called the *domain* of \mathcal{A} and is denoted $\text{dom}(\mathcal{A})$. Given partial assignments \mathcal{A} and \mathcal{A}' , we say that \mathcal{A}' *extends* \mathcal{A} if $\text{dom}(\mathcal{A}) \subseteq \text{dom}(\mathcal{A}')$ and $\mathcal{A}(p_i) = \mathcal{A}'(p_i)$ for all $p_i \in \text{dom}(\mathcal{A})$. Sometimes we refer to partial assignments simply as assignments.

Recall that a set of formulas S is *satisfiable* if there is an assignment that satisfies every formula in S . For example, the set of formulas

$$S = \{p_1 \vee p_2, \neg p_2 \vee \neg p_3, p_3 \vee p_4, \neg p_4 \vee \neg p_5, \dots\}$$

is satisfied by the assignment \mathcal{A} such that $\mathcal{A}(p_i) = 1$ if i is odd and $\mathcal{A}(p_i) = 0$ if i is even.

Theorem 1 (Compactness Theorem). *A set of formulas S is satisfiable if and only if every finite subset of S is satisfiable.*

Proof. One direction of the theorem is obvious: if S is satisfiable then every finite subset is certainly satisfiable. The non-trivial direction is the converse.

Let S be a set of formulas such that every finite subset of S is satisfiable. Say that a partial assignment \mathcal{A} is *good* if it satisfies any formula $F \in S$ that only mentions propositional variables in the domain of \mathcal{A} . We first observe that for each $n \in \mathbb{N}$ there is a partial assignment \mathcal{A} with $\text{dom}(\mathcal{A}) = \{p_1, p_2, \dots, p_n\}$ that is good. To see this, consider the subset $S' \subseteq S$ consisting of all formulas that mention only propositional variables p_1, p_2, \dots, p_n . Now S' may be an infinite set, but it only contains finitely many formulas *up to logical equivalence* since there are only finitely many formulas on propositional variables p_1, p_2, \dots, p_n up to logical equivalence (2^{2^n} formulas to be precise). Since all finite subsets of S are satisfiable

¹In the previous lectures, we used such partial assignments especially when we presented algorithms. At that time, we didn't explicitly introduce this concept. We do it here now because partiality plays an important role in the proof of the Compactness Theorem, a main topic of this lecture.

we conclude that \mathcal{S}' is satisfiable by some partial assignment \mathcal{A} with $\text{dom}(\mathcal{A}) = \{p_1, p_2, \dots, p_n\}$. By construction such an assignment is good.

The central idea of the proof is to construct a sequence of good partial assignments $\mathcal{A}_0, \mathcal{A}_1, \mathcal{A}_2, \dots$ such that $\text{dom}(\mathcal{A}_n) = \{p_1, \dots, p_n\}$ and \mathcal{A}_{n+1} extends \mathcal{A}_n for each n . We construct the \mathcal{A}_n in sequence, starting with \mathcal{A}_0 , and maintaining the following induction hypothesis:

There are infinitely many good extensions of \mathcal{A}_n such that the domain of each such extension is $\{p_1, \dots, p_k\}$ for some $k \geq n$.

For the base step we define \mathcal{A}_0 to be the assignment with empty domain. Since there is a good assignment with domain $\{p_1, \dots, p_n\}$ for every n , there are infinitely many good assignments that extend \mathcal{A}_0 ; thus \mathcal{A}_0 satisfies (*).

For the induction step, suppose that we have constructed assignments $\mathcal{A}_0, \dots, \mathcal{A}_n$ such that \mathcal{A}_n satisfies (*). Consider the two assignments $\mathcal{B}, \mathcal{B}'$ that extend \mathcal{A}_n with $\text{dom}(\mathcal{B}) = \text{dom}(\mathcal{B}') = \{p_1, p_2, \dots, p_{n+1}\}$ (say $\mathcal{B}(p_{n+1}) = 0$ and $\mathcal{B}'(p_{n+1}) = 1$.) Since any proper extension of \mathcal{A}_n is an extension of either \mathcal{B} or \mathcal{B}' , it follows that one (or both) of \mathcal{B} and \mathcal{B}' has infinitely many good extensions. Define \mathcal{A}_{n+1} to be \mathcal{B} if \mathcal{B} has infinitely many good extensions; otherwise define \mathcal{A}_{n+1} to be \mathcal{B}' . Then \mathcal{A}_{n+1} satisfies (*) by construction.

There is a unique (total) assignment \mathcal{A} that extends all the \mathcal{A}_n —it is defined by $\mathcal{A}(p_n) := \mathcal{A}_n(p_n)$ for each $n \in \mathbb{N}$. We claim that \mathcal{A} satisfies all formulas in \mathcal{S} . Indeed if $F \in \mathcal{S}$ mentions propositional variables $\{p_1, \dots, p_n\}$ then \mathcal{A}_n satisfies F . It follows that \mathcal{A} also satisfies F , since \mathcal{A} extends \mathcal{A}_n . Thus \mathcal{A} satisfies all formulas in \mathcal{S} and the proof is concluded. \square

The importance of the Compactness Theorem may be more apparent from the contrapositive formulation: *if a set of formulas \mathcal{S} is unsatisfiable then some finite subset of \mathcal{S} is already unsatisfiable*. This suggests a procedure by which we can show that an infinite set of formulas \mathcal{S} is unsatisfiable. Suppose that \mathcal{S} can be enumerated by some algorithm as

$$\mathcal{S} = \{F_1, F_2, F_3, \dots\}$$

Then for each $n \in \mathbb{N}$ we test whether the finite set $\{F_1, \dots, F_n\}$ is unsatisfiable (using, say, truth tables or some other method). The Compactness Theorem guarantees that if \mathcal{S} is not satisfiable we will detect that fact after a finite amount of time. On the other hand if \mathcal{S} is satisfiable then the above procedure will not terminate.

2 Application: Graph colouring

Let us consider an application of the compactness theorem to prove a purely combinatorial result. Recall that a graph $G = (V, E)$ is *k-colourable* if there is a function $c : V \rightarrow \{1, \dots, k\}$ mapping the set of vertices to a set of k colours such that adjacent vertices do not have the same colour, i.e., $(u, v) \in E$ implies $c(u) \neq c(v)$. Let us say that $H = (V_1, E_1)$ is a *subgraph* of G if $V_1 \subseteq V$ and $E_1 \subseteq E$.

Theorem 2. *Let $G = (V, E)$ be a graph with set of vertices $V = \{v_i : i \in \mathbb{N}\}$. Suppose that every finite subgraph of G is k -colourable. Then G is k -colourable.*

Proof. Recall how we reduced k -colouring to propositional satisfiability. Introduce propositional variables $P_{v,i}$, for each $v \in V$ and $1 \leq i \leq k$, interpreted as “vertex v has colour i ”. We consider the following propositions:

- $F_v := \bigvee_{i=1}^k P_{v,i}$ (vertex v has some colour)
- $G_v := \bigwedge_{i=1}^k \bigwedge_{j=i+1}^k (\neg P_{v,i} \vee \neg P_{v,j})$ (vertex v has at most one colour)
- $H_{u,v} := \bigwedge_{i=1}^k (\neg P_{u,i} \vee \neg P_{v,i})$ (vertices u and v don't have the same colour)

Now define $\mathcal{S} = \{F_v, G_v : v \in V\} \cup \{H_{u,v} : (u,v) \in E\}$. We claim that \mathcal{S} is satisfiable if and only if the graph G has a k -colouring. Indeed, given such a colouring c , define an assignment \mathcal{A} by $\mathcal{A}(P_{v,i}) = 1$ if and only if $c(v) = i$. Then it is clear that \mathcal{A} satisfies \mathcal{S} . Conversely, given an assignment \mathcal{A} satisfying \mathcal{S} we can define a k -colouring c by $c(v) = i$ iff $\mathcal{A}(P_{v,i}) = 1$.

By assumption, every finite subgraph of G has a k -colouring. It follows that every finite subset of \mathcal{S} is satisfiable. By the Compactness Theorem it must be that \mathcal{S} is satisfiable, and thus G itself is k -colourable. \square

3 Discussion

The proof of the Compactness Theorem bears a different character to the rest of the proofs in this course. In general, our proofs are *constructive*: a statement that something exists is proved by giving an algorithm for producing that “something”. For example, our proof that a 2-CNF formula with consistent implication graph is satisfiable consisted of an algorithm to construct a satisfying assignment given such a graph. By contrast, the proof of the Compactness Theorem does not tell us anything about the assignment that satisfies the set of formulas \mathcal{S} in the statement of theorem—merely that it exists.

Those who are studying topology might like to note that the Compactness Theorem is equivalent to the statement that the set of propositional assignments is compact under its “natural” topology. Here we identify the set of propositional assignments with the set $\{0,1\}^{\mathbb{N}}$, and endow the latter with the product topology, where $\{0,1\}$ has the discrete topology. By Tychonoff’s theorem, $\{0,1\}^{\mathbb{N}}$ is compact. Then, we can prove the equivalence of the following two statements:

- If a set of formulas \mathcal{S} is unsatisfiable, then there exists a finite subset of \mathcal{S} that is not satisfiable.
- An open cover of $\{0,1\}^{\mathbb{N}}$ with basic open sets has a finite sub-cover.