COMPLEXITY OF FINITELY PRESENTED ALGEBRAS

Dexter Kozen

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Department of Computer Science
Cornell University
Ithaca, New York 14853
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Abstract

An algebra $\mathcal{A}$ is finitely presented if there is a finite set $G$ of generator symbols, a finite set $O$ of operator symbols, and a finite set $\Gamma$ of defining relations $x=_{\mathcal{A}} y$ where $x$ and $y$ are well-formed terms over $G$ and $O$, such that $\mathcal{A}$ is isomorphic to the free algebra on $G$ and $O$ modulo the congruence induced by $\Gamma$.

The uniform word problem, the finiteness problem, the triviality problem (whether $\mathcal{A}$ is the one element algebra), and the subalgebra membership problem (whether a given element of $\mathcal{A}$ is contained in a finitely generated subalgebra of $\mathcal{A}$) for finitely presented algebras are shown to be $\leq_m^{\log}$-complete for P. The schema satisfiability problem and schema validity problem are shown to be $\leq_m^{\log}$-complete for NP and co-NP, respectively. Finally, the problem of isomorphism of finitely presented algebras is shown to be polynomial time many-one equivalent to the problem of graph isomorphism.

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1. Introduction

In this paper we study the complexity of some decision problems of finitely presented algebras, a class of simple algebraic structures.

An algebra \( \mathcal{A} \) is finitely presented if there is a finite set \( G \) of generator symbols, a finite set \( O \) of operator symbols of various finite arities, and a finite set \( \Gamma \) of axioms or defining relations of the form \( x \equiv y \), where \( x \) and \( y \) are well-formed terms over \( G \) and \( O \), such that \( \mathcal{A} \) is isomorphic to the free algebra on \( G \) and \( O \) modulo the congruence induced by \( \Gamma \). That is, if \( \tau \) is the free algebra (algebra of terms) over \( G \) and \( O \), and \( \equiv \) is the smallest congruence relation satisfying the relations \( \Gamma \), then \( \mathcal{A} \) is isomorphic to the quotient algebra \( \tau / \equiv \) with domain \( \{ [x] \mid x \in \tau, [x] \text{ is the } \equiv \text{-congruence class of } x \} \).

For example, the two element Boolean algebra is presented by

\[
G = \{0, 1\} \\
O = \{\land, \lor, \neg\} \\
\Gamma = \{0 \land 0 \equiv 0, 0 \land 1 \equiv 0, 1 \land 0 \equiv 1, 1 \land 1 \equiv 1, \\
0 \lor 0 \equiv 0, 0 \lor 1 \equiv 1, 1 \lor 0 \equiv 1, 1 \lor 1 \equiv 1, \\
\neg 0 \equiv 1, \neg 1 \equiv 0\}.
\]

All algebras with finite domains are finitely presented. Finitely presented algebras may be infinite, but infinite groups and semigroups are never finitely presented, since an axiom schema (a rule representing infinitely many axioms) is needed to postulate associativity.
There is a strong relationship between finitely presented algebras and the finite tree automata of Thatcher and Wright\textsuperscript{12} and Doner\textsuperscript{13}. This relationship is summed up in the following theorem, analogous to a theorem of Nerode\textsuperscript{9} regarding the representation of regular sets over a semigroup.

**Theorem**

L is a regular tree language (accepted by a finite tree automaton) over τ iff L is a union of congruence classes of a finitely generated congruence relation on τ of finite index. ■

Moreover, all congruence classes of any finitely generated congruence, finite index or not, are regular tree languages. The set of terms representing a single element of a finitely presented algebra is such a class.

Finite tree automata appear in diverse settings. Not only do they have a substantial theory of their own (see [10,11] for a good bibliography), but they have also been used in logic to show the decidability of some second order theories\textsuperscript{12,13,14} and in formal language theory to study derivation trees of context free grammars (see [10,11]). In view of the above theorem, the complexity results presented here should apply to those areas.

Most of the decision problems addressed herein can be restated as problems of tree replacement systems, hence our complexity results carry over into that area.

Finally, very recent results, notably [1,2], have pegged down the complexities of various decision problems in different
algebras. The present results fill a large gap here, and so would be essential to a general theory of the complexity of algebraic decision problems.

In spite of the above, the results presented here are most interesting not for any of these reasons. Their real interest lies in the generality and expressive power of the language of universal algebra. The finite structures that interest computer scientists, e.g. graphs, are easily represented as finitely presented algebras, and many known complete problems for P, NP, etc., can be reformulated easily as natural questions about finitely presented algebras, as evidenced by the trivial (often gsm) reductions from known complete problems to the problems discussed in this paper. Thus finitely presented algebras should be viewed as a unifying framework in which many of the interesting questions of low-level complexity can be reformulated.

In §3 we give several natural problems of finitely presented algebras which are $\leq^m_{\log}$-complete for P. These problems generalize known problems complete for P. In §4 we look at axiom schemata of the form $x \equiv y$ where $x$ and $y$ are terms with variables, and show that the schema satisfiability problem is $\leq^m_{\log}$-complete for NP and the schema validity problem is $\leq^m_{\log}$-complete for co-NP. In §5 we show that the problem of isomorphism of finitely presented algebras is polynomial time many-one equivalent to the problem of graph isomorphism.
2. **Preliminaries**

**Definition**

Let $\langle M, \text{ARITY} \rangle$ be a ranked alphabet, i.e., $M$ is a finite set of symbols and $\text{ARITY}: M \to N$, where $N$ is the set of nonnegative integers. Partition $M$ into two sets:

$$G = \{ a \in M \mid \text{ARITY}(a) = 0 \} \text{ are generator symbols,}$$

$$O = \{ a \in M \mid \text{ARITY}(a) > 0 \} \text{ are the operator symbols.}$$

We will use variables $a, b$ to denote elements of $G$ and $0, 0$ to denote elements of $O$. Let $M^*$ be the set of finite length strings over $M$.

**Definition**

The set of **terms** over $M$ is the smallest subset of $M^*$ such that

i) all elements of $G$ are terms;

ii) if $0$ is $m$-ary (i.e. $\text{ARITY}(0) = m$) and $x_1, \ldots, x_m$ are terms, then $0x_1 \ldots x_m$ is a term.

Denote the set of terms by $\tau$. Variables $w, x, y, z$ will range over terms.

$\tau$ may be viewed as the domain of an algebra with operations $0$ defined by

$$0(x_1, \ldots, x_m) = 0x_1 \ldots x_m \text{ if } 0 \text{ is } m\text {-ary.}$$
In this light we will refer to \( \mathcal{T} \) as the **free algebra** on \( M \).

**Definition**

\( x \triangleleft y \) if \( x \) is a (not necessarily proper) subterm of \( y \).

\( x[y/z] \) is the term \( x \) with all occurrences of the term \( y \) in \( x \) replaced by \( z \).

If \( y' \) is a particular occurrence of a term \( y \) as a subterm of \( x \), then \( x[y'/z] \) is the term \( x \) with that occurrence only replaced by \( z \).

**Definition**

A binary relation \( \simeq \) on \( \mathcal{T} \) is a **congruence** provided

i) \( \simeq \) is an equivalence relation

ii) if \( \theta \) is \( m \)-ary and \( x_1, \ldots, x_m, y_1, \ldots, y_m \) are terms such that \( x_i \simeq y_i \), \( i \in \text{isim} \), then

\( \theta x_1 \ldots x_m \simeq \theta y_1 \ldots y_m \). 

In the above definition, ii) guarantees that the operations \( \theta \) are well-defined on \( \simeq \)-congruence classes, thus we can form the quotient algebra \( \mathcal{T}/\simeq \) with domain

\( \{[x] \mid x \in \mathcal{T}, [x] \text{ is the } \simeq \text{-congruence class of } x \} \)

and operations \( \theta \).

**Definition**

Let \( \Gamma \) be a set of unordered pairs of terms. These pairs will be written \( x \triangleleft y \) and will be called **axioms** or **defining relations**. Define \( \Xi_\Gamma \) to be the smallest congruence on \( \mathcal{T} \) satisfying the axioms of \( \Gamma \), and let
\[ [x]_\Gamma = \{ y \in \tau \mid x \equiv \Gamma y \}. \]

We will omit the subscript \( \Gamma \) from \( \equiv \) and \( [x]_\Gamma \) when it is understood. It is straightforward to show that \( x \equiv \Gamma y \) iff it can be deduced from:

i) \( x \equiv \Gamma x \)

ii) \( x \equiv \Gamma y, y \equiv \Gamma x \)

iii) \( x \equiv \Gamma y, y \equiv \Gamma z \)

\[ \frac{}{x \equiv \Gamma z} \]

iv) \( x_1 \equiv \Gamma y_1, \ldots, x_m \equiv \Gamma y_m, \text{ ARITY}(\theta) = m \)

\[ \theta x_1 \ldots x_m \equiv \Gamma \theta y_1 \ldots y_m \]

v) \( x \equiv \Gamma y \) for all axioms \( x \equiv y \) of \( \Gamma \).

**Definition**

An algebra \( \mathcal{A} \) is presented by \( \langle M, \text{ARITY}, \Gamma \rangle \) if \( \mathcal{A} \) is (isomorphic to) \( \Gamma / \equiv \). The triple \( \langle M, \text{ARITY}, \Gamma \rangle \) is called a presentation of \( \mathcal{A} \). \( \mathcal{A} \) is finitely presented if a presentation can be found with \( \Gamma \) a finite set.

It is convenient to represent terms as labeled trees, as follows:

i) if \( a \in G \) then \( a \) is represented by a single vertex with label \( a \).

ii) \( \theta y_1 \ldots y_m \) is represented by

\[ \text{Diagram:} \]

- A root node labeled \( \theta \) with three children labeled 1, 2, and \( m \).
- Each child node has a label, with \( y_1, y_2, \ldots, y_m \) as children of the terminal nodes.
where the root has label \( \emptyset \) and \( y_i \) is the tree representation of term \( y_i \).

This representation has three immediate advantages:

1) We can give a presentation consisting of a finite set of trees labeled as above, with an extra undirected edge set AXIOM such that, if \( y \) is the tree representation of term \( y \), \( x \equiv y \) is an axiom \( \leftrightarrow \) the roots of \( x \) and \( y \) are connected with an AXIOM edge.

We no longer need to specify \( M \) and ARITY, since these are implicit in the new representation.

2) We can represent terms with common subterms more concisely, by "factoring out" the common subterm, i.e. representing a set of trees as a dag.

E.g. terms \( \emptyset \text{abc}, \emptyset \emptyset \text{abc}\emptyset \text{abca} \)

may be represented as trees by

\[ \text{Diagram} \]

\[ \text{Another Diagram} \]
and then as dags, after factoring, by

3) We have a conceptually simpler deductive system for proving congruence of terms:

Define $x \rightarrow y$ (x derives y in one step) if there is an axiom $z \equiv w$ in $\Gamma$ and an occurrence of $z$ as a subtree of $x$ such that if that occurrence of $z$ is replaced by $w$, then the result is $y$.

Let $\Rightarrow$ be the reflexive transitive closure of $\rightarrow$. The following is proved by an easy induction:

**Theorem**

$x \equiv y \iff x \Rightarrow y$.

For these reasons we will henceforth adopt the new representation, and use the words "term" and "tree" interchangeably. We will allow trees to be represented as dags by factoring out common subterms, and we will consider a presentation to be given by a dag with AXIOM edges, as outlined above. We will reuse the symbol $\Gamma$ for the dag representing presentation $<M, ARITY, \Gamma>$. 
Definition

A proof of \( x = y \) is a sequence \( x_1 \ldots x_n \) of terms such that \( x = x_1 + x_2 + \ldots + x_n = y \).

The root of the transformation in \( x \rightarrow y \) is the root of the subtree replaced. We say the entire tree is replaced in \( x \rightarrow y \) if the subtree of \( x \) that is replaced is the whole tree \( x \).

Definition

Let \( \Gamma \) be given. The following sets will be used throughout:

\[
R_\Gamma = \{ x \mid \text{there is an axiom } y \leq z \text{ in } \Gamma \text{ and } x \leq y \}.
\]

\[
r_\Gamma = \{ [x]_\Gamma \mid x \in R_\Gamma \}.
\]

Thus \( R_\Gamma \) is the set of terms appearing in the presentation, and \( r_\Gamma \) are the elements of the presented algebra represented by terms in \( R_\Gamma \). The subscript \( \Gamma \) will be omitted when understood.

It is assumed the reader is familiar with the complexity classes \( P, NP, \) etc., the reducibilities \( \leq^m \log \) and \( \leq^m \), and the notion of completeness.
3. Problems complete for P

The obvious first task is to determine the complexity of deciding whether two terms represent the same element of the algebra.

Definition

The word problem is the set

\[ WP = \{ <\Gamma, x, y> \mid x \equiv_\Gamma y \}. \]

WP would more accurately be called the uniform word problem, since the presentation \( \Gamma \) is an input parameter.

Theorem 1

\( WP \in P. \)

Proof

Let \( <\Gamma, x, y> \) be input. Let \( \Gamma^+ \) be the graph \( \Gamma \) plus the vertices and edges of \( x \) and \( y \). Let \( R^+ = R_{\Gamma^+}^+ \). We will describe a polynomial time algorithm to construct a new undirected edge set \( E \) on \( \Gamma^+ \) so that \( \forall z, w : z \equiv R^+ w \), the roots of \( z \) and \( w \) are connected by an \( E \) edge (henceforth \( z \equiv w Ew \)) iff \( z \equiv w \).

**Step 0.** Add edges \( wEw, w \in R^+ \).

Add edges \( wEz \) for all axioms \( w \equiv a \).

**Step n.** If \( u, v, w \in R^+ \) and \( uEw, vEw \), then add edge \( uEw \). If \( \emptyset y_1 \ldots y_m, \emptyset x_1 \ldots x_m \in R^+ \) and \( x_1 Ey_1 \), \( 1 \text{ is a } \text{ variable, then add edge } \emptyset x_1 \ldots x_m E\emptyset y_1 \ldots y_m \). If no new edges were added at Step \( n \), then stop.
The algorithm is clearly polynomial, since at most $n^2$ edges can be added.

Claim

$$\forall w, z \in R^+ \; wEz \iff w \equiv z.$$  

Proof of claim

(→) Easy induction on the step at which the edge was added, since no edge is added unless forced to be by the properties of $\equiv$.

(←) Define a relation $\mathcal{C}$ on congruent pairs of terms, as follows:

$$<x, y> \in \mathcal{C} <z, w> \iff \text{either}$$

i) the shortest proof $x \overset{*}{\Rightarrow} y$ is shorter than

the shortest proof $z \overset{*}{\Rightarrow} w$; or

ii) the shortest proofs $x \overset{*}{\Rightarrow} y$ and $z \overset{*}{\Rightarrow} w$ are the same

length and $x \overset{*}{\Rightarrow} z$, $y \overset{*}{\Rightarrow} w$.

Clearly $\mathcal{C}$ is a well-founded relation, so we proceed by induction on $\mathcal{C}$.

Let $x, y \in R^+.$

Basis: length of proof $x \overset{*}{\Rightarrow} y = 0$.

Then $x = y$ and $xEy$ at step 0.

Induction step: $x \overset{*}{\Rightarrow} y$ is a nonzero-length shortest proof.

One of the following two cases must hold:

Case 1: The entire tree is replaced somewhere in $x \overset{*}{\Rightarrow} y$.

In this case $3z, w \overset{*}{\Rightarrow} z$, $z \equiv w$ is an axiom,

$w \overset{*}{\Rightarrow} y$. But $z, w \in R^+$ hence $xEz$ and $wEy$ by

induction hypothesis, since they are congruent via shorter proofs; and $zEw$ in step 0. Thus $xEy$ within two more steps.
Case 2: The entire tree is never replaced in $x^*y$.
Then $x = 0x_1...x_m$, $y = 0y_1...y_m$, and $x_i^*y_i$ via
a proof of length shorter than or equal to $x^*y$
(the proof $x_i^*y_i$ is given by the transformations
applied to the interior of $x_i$ in the proof $x^*y$)
and $x_i \not\in x$, $y_i \not\in y$, lsism, hence $<x_i, y_i> \in$
$<x, y>$, lsism. But all $x_i, y_i \in R^+$, hence by
the induction hypothesis, $x_iEy_i$, lsism. Then
in the next step of the algorithm, $xEy$.

Definition

An instance of the circuit value problem (CVP) is a
list $B$ of assignments to variables $C_1, C_2, ..., C_n$ of the
form

$$C_i = 0,$$
$$C_i = 1,$$
$$C_i = C_j \wedge C_k, j, k < i,$$
$$or C_i = C_j \vee C_k, j, k < i,$$

such that each $C_i$ appears on the left side of an assignment
exactly once. $B$ is in CVP provided $\text{val}(C_n) = 1$, where $\text{val}(C_i)$
is the Boolean value of $C_i$ computed from the list of assignments
in the obvious way.

As demonstrated by Ladner, CVP is $\leq_{\text{log}}^m$-complete for $P$.

Theorem 2

$\text{CVP} \leq_{\text{log}}^m \text{WP}$. 
Proof

Given the above instance of CVP, take

\[ G = \{C_1, \ldots, C_n, 0, 1\} \]
\[ O = \{\wedge, \vee\} \]
\[ \Gamma = B \cup \{0 \vee 0 \equiv 0, 0 \vee 1 \equiv 1, 1 \vee 0 \equiv 1, 1 \vee 1 \equiv 1, \]
\[ 0 \wedge 0 \equiv 0, 0 \wedge 1 \equiv 0, 1 \wedge 0 \equiv 0, 1 \wedge 1 \equiv 1\}. \]

The restrictions on \( B \) and the eight extra axioms guarantee that the algebra presented by \( \Gamma \) is the two element lattice, and \( B \in \text{CVP} \) iff \( C_n \equiv 1 \) iff \( \langle \Gamma, C_n, 1 \rangle \in \text{WP} \). \(\square\)

Observe that CVP is really a special case of the word problem, as shown by the trivial (gsm) reduction.

Corollary 3

WP is \( m_{\log} \)-complete for P. \(\square\)

We now wish to show the following three problems complete for P.

Definition

\[ \text{TRIV} = \{\Gamma \mid \Gamma \text{ presents the trivial (one element) algebra}\}. \]
\[ \text{FIN} = \{\Gamma \mid \Gamma \text{ presents a finite algebra}\}. \]
\[ \text{GEN} = \{\langle \Gamma, x_1, \ldots, x_n, y \rangle \mid [y] \text{ is contained in the subalgebra of } \Gamma/\Xi \text{ generated by } [x_1], \ldots, [x_n]\}. \]

Theorem 4

\[ \text{TRIV} \in \text{P}. \]
Proof

Use the algorithm of Theorem 1 to decide for all $a, b \in G$ whether $a \equiv b$, then for all $\emptyset \in \mathcal{O}$ whether $\emptyset a \ldots a \equiv a$.

Theorem 5

$$CVP \leq^m_{log} \text{TRIV}.$$  

Proof

Let $B$ be an instance of $CVP$. Construct $\Gamma$ as in Theorem 2 and let $\Gamma' = \Gamma \cup \{c_n \equiv 0\}$. Then

$B \in CVP$ iff $C_n \equiv_\Gamma 1$ iff $1 \equiv_\Gamma 0$ iff $\Gamma'/\equiv_\Gamma$ is trivial.

Corollary 6

$\text{TRIV}$ is complete for P.

We now wish to show $\text{GEN}$ is complete for P. $\text{GEN}$ is a more general formulation of the problem of the same name of Jones and Laaser.

Theorem 7

$\text{GEN} \in P$.

Proof

Given $<\Gamma, x_1, \ldots, x_n, y>$, let $\mathcal{A}$ be the subalgebra of $\Gamma/\equiv$ generated by $[x_1], \ldots, [x_n]$, and let $\mathcal{E}$ be the subalgebra of $\Gamma$ generated by $x_1, \ldots, x_n$. Then $\mathcal{E}/\equiv$ is isomorphic to $\mathcal{A}$, and $[y] \in \mathcal{A}$ iff $\exists x \in \mathcal{E}$ such that $y \equiv x$.

Let $\Gamma^+ = \Gamma \cup \{x_1, \ldots, x_n, y\}$ and let $R^+ = R^+_{\Gamma^+}$. Consider the following algorithm to mark elements of $R^+$ (vertices of $\Gamma^+$):
Step 0. Run the algorithm of Theorem 1 on $R^+$ to determine for all $x,w \in R^+$ whether $x \equiv w$. Mark each $x_i$.

Step n. If $\emptyset y_1 \ldots y_m \in R^+$ and $y_1, \ldots, y_m$ are marked, mark $\emptyset y_1 \ldots y_m$.

If $x,w \in R^+$, $x \equiv w$, and $x$ is marked, mark $w$.

If no new terms are marked, stop.

The algorithm is clearly polynomial. The following claim establishes the result.

Claim

$[y] \in \mathcal{A}$ iff $y$ is marked by the above algorithm.

Proof of claim

$(\rightarrow)$ clear.

$(\leftarrow)$ let $C_0 = \{x_1, \ldots, x_n\}$

$$C_{k+1} = \{\emptyset y_1 \ldots y_m \mid y_1, \ldots, y_m \in C_k\} \cup C_k$$

Then $\nu_k C_k = L$. Let $y \in R^+$ such that $[y] \in \mathcal{A}$. Then $\exists x \in L y \equiv x$.

We prove the result by induction on the least $k$ such that $\exists x \in C_k y \equiv x$.

Basis: $k=0$. Then $y \equiv x_1$, and $x_1$ is marked at step 0, hence $y$ is marked at step 1.

Induction step: $k>0$. Then $y \equiv \emptyset y_1 \ldots y_m$ and $y_1, \ldots, y_m \in C_{k-1}$.

We have no guarantee that $y_1, \ldots, y_m \in R^+$ however, hence we cannot claim that $y_1, \ldots, y_m$ are marked. But let us consider a shortest proof $y \vdash \emptyset y_1 \ldots y_m$. One of the following two cases occurs:
Case 1: The entire tree is never replaced in $y \ast \theta y_1 \ldots y_m$. Then $y = \theta z_1 \ldots z_m$ and each $z_i \in y_i$. But each $y_i \in C_{k-1}$ hence $[z_i] \in A$, thus since $z_i \in R^+$, by the induction hypothesis $z_i$ is eventually marked. Then when each $z_i$ is marked, $y$ is marked in the next step.

Case 2: $y \ast z$, $z \equiv w$ is an axiom, $w = \theta w_1 \ldots w_m$ and $w_i \equiv y_i$, $1 \leq i \leq m$. I.e., $z \ast w$ is the last time in the proof $y \ast \theta y_1 \ldots y_m$ that the entire tree is replaced. Then $w_i \in R^+$ and $y_i \in C_{k-1}$, hence by the induction hypothesis, $w_i$ is eventually marked, and $w \in R^+$ hence $w$ is marked; but $y \equiv w$, thus $y$ is marked in the next step.

**Corollary 8**

GEN is complete for $P$.

**Proof**

There is a trivial reduction from Jones' and Laaser's GEN to our GEN. The details are left to the reader.

We now turn to the finiteness problem. Let $\Gamma$ be a presentation of $A$.

**Lemma 9**

$A$ is finite iff $A = r_\Gamma$.

**Proof**

($\ast$) Trivial.

($\ast$) Clearly $r_\Gamma \subseteq A$. Now assume there is an $x$ such that
[x] \in \mathcal{A} - R_{\tau}, \text{ i.e. } \forall y \in R_{\tau} \ x \not\equiv y. \text{ Let } x \text{ be } \preceq \text{-} \text{minimal in its congruence class. Define a set of terms }$
\begin{align*}
x_0 &= x \\
x_{n+1} &= o x_n \ldots x_n
\end{align*}$
where \( o \) is any operator. For all \( j > 0 \), if \( x_0 \equiv x_j \), then \( x \not\equiv x_j \) and \( x_j \not\equiv x \). But in the proof \( x_j \not\equiv x \), no ancestor of \( x \) in \( x_j \) can ever be the root of a transformation, else \( x \) is congruent to a term in \( R_{\tau} \), hence \( x \not\equiv y \) in the proof \( x_j \not\equiv x \), where \( y \) is a proper subterm of \( x \), contradicting the assumption of \( \preceq \text{-} \text{minimality}. \)

Proceeding by induction, if \( x_{i+1} \equiv x_{j+1} \), \( i \neq j \), then \( o x_i \ldots x_i \not\equiv o x_j \ldots x_j \). But since \( x \not\equiv x_i \) and no transformation can be applied to an ancestor of \( x \) in \( x_i \), it must be that \( x_i \equiv x_j \), contradicting the induction hypothesis.

Hence \( \{[x_0],[x_1],\ldots\} \) is an infinite subset of \( \mathcal{A} \), contradicting the finiteness of \( \mathcal{A} \).

The above lemma uncovers a surprising fact: a finite algebra can be no bigger than its smallest finite presentation. Thus to give a finite presentation of a finite group, say, we might as well write down its multiplication table.

We wish to show a ptime algorithm to decide finiteness. The straightforward approach would be to show that if \( \mathcal{A} \) is infinite then there is a small term (i.e. one of the form \( o y_1 \ldots y_m \) where \( y_1, \ldots, y_m \in R_{\tau} \)) with \( [o y_1 \ldots y_m] \not\in R_{\tau} \), thus we would need only to run the algorithm of Theorem 1 on each small term \( y \) and each
x ∈ R_{Γ} to see if y ≡ x. However, the arity of the operators is an input parameter and can grow as much as linearly with the size of the presentation, so there may be too many $y_1 \ldots y_m$ to try; hence this approach works only when the arity of the operators is bounded. We circumvent this problem with the following construction.

**Definition**

The **characteristic graph** of a presentation Γ, denoted $χ_{Γ}$, is a labeled directed graph with vertex labels 0 and edge labels \{1, 2, \ldots, k\} where k is the maximum arity of any operator in 0. The vertex set of $χ_{Γ}$ consists primarily of unlabeled vertices $x_{Γ}$, plus other labeled vertices and labeled and unlabeled edges such that $χ_{Γ}$ is the smallest graph in which

\[
[0y_1 \ldots y_m]
\]

\[
\begin{array}{c}
\text{1} \\
\text{[y_1]} \\
\text{2} \\
\text{[y_2]} \\
\text{m} \\
\text{[y_m]} \\
\end{array}
\]

appears as a subgraph of $χ_{Γ}$ for every $0y_1 \ldots y_m \in R_{Γ}$.

**Example**

The $Γ$ in Theorem 2 presents the two element lattice. Its characteristic graph is:
\( \chi_\Gamma \) is meant to represent the interaction of the elements of \( \Gamma \) under the operations \( \circ \). \( \chi_\Gamma \) is constructible in polynomial time: we can just run the algorithm of Theorem 1 to determine \( \Xi_\Gamma \)-classes (they appear as cliques of \( E \)-edges), then for each \( \emptyset y_1 \ldots y_m \) in \( \Gamma \), add a vertex \( \emptyset \) and edges

\[
[ \emptyset y_1 \ldots y_m ]
\]

```
  \emptyset
  \downarrow
  1 \quad m
  \downarrow
[y_1] \ldots [y_m].
```

Lemma 10

For \( x, y_1, \ldots, y_m \in \Gamma \), \( x \equiv \emptyset y_1 \ldots y_m \) iff
appears in $\chi_r$.

Proof

$(\rightarrow)$ If the subgraph pictured appears in $\chi_r$, then by construction of $\chi_r$ it must be that $z_1, \ldots, z_m \equiv z_1 \ldots z_m$ and $z_i \equiv y_i$, $l \equiv m$. Then $x \equiv z_1 \ldots z_m \equiv y_1 \ldots y_m$.

$(\rightarrow)$ Consider a proof $x \vdash y_1 \ldots y_m$. If the entire tree is ever replaced, $\exists z_1 \ldots z_m \ x \vdash w \vdash z_1 \ldots z_m$ and $z_i \vdash y_i$, $l \equiv m$, and $w \equiv z_1 \ldots z_m$ is an axiom. If the entire tree is never replaced, then $\exists z_1 \ldots z_m \ x \equiv z_1 \ldots z_m \vdash y_1 \ldots y_m$ and $z_i \vdash y_i$, $l \equiv m$. In either case,

appears in $\chi_r$, and $[\equiv z_1 \ldots z_m] = [x]$ and $[z_i] = [y_i]$, $l \equiv m$. □
Lemma 11

$\mathcal{A}$ is finite iff for all $m$, for all $m$-ary $\theta$ and $[y_1], \ldots, [y_m] \in R_\Gamma$, there is an $[x] \in R_\Gamma$ such that

\[ [x] \]

\[ \rightarrow \]

\[ 1 \]

\[ \rightarrow \]

\[ m \]

\[ [y_1] \ldots [y_m] \]

appears in $x_\Gamma$.

Proof

By the previous lemma, for all $\theta$ and $[y_1], \ldots, [y_m]$ there is an $[x] \in R_\Gamma$ such that the above graph appears in $x_\Gamma$ iff $R_\Gamma$ is closed under all the operations $\theta$. But since $\mathcal{A}$ is generated by $\{ [a] \mid a \in G \} \subseteq R_\Gamma$, this occurs iff $\mathcal{A} = R_\Gamma$. By Lemma 9, this occurs iff $\mathcal{A}$ is finite.

Theorem 12

$FIN \in P$.

Proof

Construct $x_\Gamma$, and for each $m$-ary $\theta$, cycle through all $[y_1], \ldots, [y_m]$, rejecting if we ever find a $[y_1], \ldots, [y_m]$ such that for no $x \in R_\Gamma$ does
Theorem 13

\[ \text{CVP} \leq_{\log} \text{FIN}. \]

Proof

We use the presentation $\Gamma'$ constructed in Theorem 5, such that $\Gamma'$ presents either the trivial algebra or the two element lattice, and

\[ \Gamma/\Xi_{\Gamma'} \text{ is trivial } \iff \mathcal{B} \in \text{CVP}. \]

Append another generator symbol $b$ to $G$, and the axioms

\{ $b\wedge b = 0$, $b\wedge 0 = 0$, $0\wedge b = 0$, $b\vee b = 0$, $b\vee 0 = 0$, $0\vee b = 0$ \} to $\Gamma'$ to get $\Gamma''$. It is left to the reader to verify that $\Gamma''$ presents the trivial algebra if $\Gamma'$ does, an infinite algebra otherwise.

Corollary 14

FIN is complete for P.
4. Two schema problems complete for NP and co-NP

Definition

Let $\tau = \{\text{terms over } G \text{ and } 0\}$, as usual.

Let $V = \{v_1, \ldots, v_m\}$ be a set of variable symbols,

$G^+ = G \cup V,$

$\tau^+ = \{\text{terms over } G^+ \text{ and } 0\}.$

Thus $\tau^+$ is the set of terms with occurrences of variables $V$.

An assignment to variables is a map $I: V \to \tau$. If we take $I(a) = a$ for a $\in G$, then $I$ extends uniquely to a homomorphism $\tau^+ \to \tau$, which we will also denote by $I$.

A schema is a formula $x \equiv y$, where $x, y \in \tau^+$. Given $\Gamma$, a schema $x \equiv y$ is satisfiable in $\tau/\equiv_\Gamma$ if there is an assignment $I$ such that $I(x) \equiv_\Gamma I(y)$. I.e., $x \equiv y$ is satisfiable if there is an interpretation of terms with variables over $\tau/\equiv_\Gamma$ such that $x$ and $y$ represent the same element of $\tau/\equiv_\Gamma$.

A schema $x \equiv y$ is valid in $\tau/\equiv_\Gamma$ if for all assignments $I$, $I(x) \equiv_\Gamma I(y)$.

Definition

The schema satisfiability problem is the set

$\text{SATIS} = \{< \Gamma, x, y> \mid \text{schema } x \equiv y \text{ is satisfiable}\}.$

The schema validity problem is the set

$\text{VALID} = \{< \Gamma, x, y> \mid \text{schema } x \equiv y \text{ is valid}\}.$
Observe that the Boolean satisfiability problem of Cook\textsuperscript{5} is a special case of SATIS, where the algebra presented by $\Gamma$ is the two element Boolean algebra. It will be left to the reader to verify that there are very easy (gsm) reductions from the Boolean satisfiability problem to SATIS and from the Boolean tautology problem to VALID.

Fix $< \Gamma, x, y >$. $V = \{v_1, \ldots, v_m\}$ are the variables occurring in $x$ and $y$. Define $\Gamma^+$ to be the graph $\Gamma$ plus $x$ and $y$, and let $R^+ = R_{\Gamma^+}$.

In the following, if $X = \{v_{i_1}, \ldots, v_{i_k}\}$ is a set of variables, and $I: X \rightarrow \tau$ is a partial assignment, we will write

$$z[X \setminus I(X)]$$

in place of

$$z[v_{i_1} \setminus I(v_{i_1}), \ldots, v_{i_k} \setminus I(v_{i_k})].$$

Definition

An assignment $I: V \rightarrow \tau$ is well-specified on $X \subseteq V$ if $gf: X \rightarrow R^+$ such that if $v \in X$ then

$$I(v) = f(v)[X \setminus I(X)].$$

I.e. the values of $I$ on members of $X$ are uniquely determined by assigning terms in $R^+$ to variables. We wish to show that if $x \equiv y$ is satisfiable, then it is satisfiable via an assignment $I$ well-specified on $V$. This will allow us to guess the map $f$ and use the word problem algorithm to verify that $I(x) \equiv I(y)$, thereby showing SATIS $\in$ NP.
Lemma 15

If $x \equiv y$ is satisfiable, then it is satisfiable via an assignment well-specified on $V$.

Proof

Suppose $x \equiv y$ is satisfiable, and $X \not= V$ is a maximal subset of $V$ such that there is an assignment $I$ well-specified on $X$ satisfying $x \equiv y$.

Let us call a term $z \in \mathcal{T}$ well-specified if it is congruent to some $w[X \setminus I[X]]$, where $w \in R^+$; ill-specified otherwise. We observe that if a term $w \in R^+$ has variables exclusively in $X$, then any subterm of $w[X \setminus I[X]]$ is well-specified. In particular, any term in $R^+$ without variables is well-specified, and every $I(v)$ for $v \in X$ is well specified.

Define a binary relation $< \not= $ on ill-specified terms by: $w < z$ iff $w$ is congruent to a proper subterm of $z$. The fact that $<$ is defined only on ill-specified terms guarantees transitivity. For, suppose $w_1 < w_2 < w_3$. Then $w_1 \equiv w_1 \not= w_2$, and $w_2 \equiv w_2 \not= w_3$, as shown.

But in a proof $w_2 \not= w_2$, no transformation can be applied to an ancestor of $w_1$, else $w_1$ would be congruent to a term in $R$, hence
If every variable is \( \alpha \)-congruent to a term in \( t \), then

\[
\{ \forall \theta \in \Delta \mid (\forall \phi \in \Delta) \psi \in \Phi \} \cup L = \Phi
\]

such that \( \psi \) is a map \( \forall \phi \in \Phi \). If \( \psi \) is satisfiable in \( \Phi \), then there is a map \( \forall \phi \in \Phi \). Thus

\[
Lem \text{ma 16}
\]

Let \( \psi = \text{true} \) and let \( x \in \Phi \) and let \( x \in x \) be a schema with variables \( \Delta \).

\[
\theta \neq \text{false} \]

contradicting the maximality of \( I \). Hence \( I \) is well-specified on \( \Delta \) and \( I \) is satisfiable on \( \Delta \). Then \( \Delta \) is satisfiable \( I \), and \( I \) is satisfiable \( I \), so \( I \) is satisfiable \( I \) and \( I \) is satisfiable \( I \).

\[
[[\forall \theta \in \Delta \mid (\forall \phi \in \Phi) \psi \in \Phi \} \cup L = \Phi
\]

\[
\forall x \neq (\forall \phi \in \Phi) \psi \in \Phi \}
\]

\[
[[\forall \theta \in \Delta \mid (\forall \phi \in \Phi) \psi \in \Phi \} \cup L = \Phi
\]

\[
\forall x \neq (\forall \phi \in \Phi) \psi \in \Phi \}
\]

\[
\forall x \neq (\forall \phi \in \Phi) \psi \in \Phi \}
\]

For all \( x \neq (\forall \phi \in \Phi) \psi \in \Phi \}

Thus

-27-
each occurrence of a subterm of \( x \) or \( y \) congruent to that


\( (\forall x)(\forall y) \) hence


occurrence of \( (\forall x) \) hence


and vice-versa. Thus there is a proof \( x + y \) "learning each

go to occurrences of \( (\forall x)(\forall y) \) in \( \ldots \) in the corresponding positions,

\( (\forall x)(\forall y) \) was well-specified. Thus all occurrences of \( x \) \( (\forall x)(\forall y) \) are

equivalent can ever be applied to an ancestor of \( (\forall x)(\forall y) \), else

let \( x \) and \( y \) be formed from \( x \) and \( y \) in this way. Then

congruent to \( (\forall x)(\forall y) \) with \( x \) and \( y \) a well-defined notion,

are \( x \) incomparsible; thus replacing all occurrences of terms

\( (\forall x)(\forall y) \) of \( x \) congruent to, we have that any

\[ [(\forall x)(\forall y)] x = x \]

\[ [(\forall x)(\forall y)] x = x \]

let \( x \in \ldots \) such that \( (\forall x)(\forall y) \) is well-founded.

Thus \( x \) restricted to \( (\forall x)(\forall y) \) is well-founded.

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Thus \( x \) restricted to \( (\forall x)(\forall y) \) is well-founded.
Proof

(+) Let $X_0 = \{ v \in V \mid f(v) \in \tau \}$. $X_0$ must be nonempty, else i) is contradicted. Similarly if $X_n$ has been defined and $X_n \subseteq V$, then $X_{n+1} = \{ v \in V \mid f(v) \text{ contains only variables in } X_n \}$ must contain a variable not in $X_n$. Thus $\nu X_n = V$. Let $p(v) = \text{least } n \text{ such that } v \in X_n$. Define $I(v) = f(v)$ for $v \in X_0$, and for $v \in X_{n+1} - X_n$ define $I(v) = f(v)[X_n \setminus I[X_n]]$. An easy induction on $p(v)$ gives $v \equiv I(v)$ for all $v \in V$.

Now we claim that $I$ satisfies $x \equiv y$. Let

$$x' = x[V \setminus I[V]],$$

$$y' = y[V \setminus I[V]].$$

We have $x' \equiv y'$, by ii). We need to show $x' \equiv y'$. Consider a proof $x' \vdash y'$. If no variable ever appears in the proof, then done. Otherwise let $n$ be the largest number such that $v \in V$ appears in the proof with $p(v) = n$. $v$ appears only via application of the axiom $v \equiv f(v)$. But the only way an ancestor of $v$ can be the root of a subsequent transformation is if a variable with a larger $p$ appears. But $v$ eventually disappears, since $y'$ has no variables, and this can occur only if $v \equiv f(v)$ is applied. Thus we can eliminate the two transformations $f(v) \vdash v$ and $v \vdash f(v)$. This process may be repeated until no variables appear in the proof $x' \vdash y'$; then $x \equiv y$.

(+) By the previous lemma, there is an assignment $I$ well-specified on $V$ satisfying $x \equiv y$, i.e. $\exists f: V \to R^+ \text{ such that } I(v) = f(v)[V \setminus I[V]]$. It is left to the reader to verify that
if $\Gamma^+ = \Gamma \cup \{ v \equiv f(v) \mid v \in V \}$ then

i) $v \equiv_{\Gamma^+} I(v)$, and

ii) $x \equiv_{\Gamma^+} y$.

Theorem 17

SATIS $\in$ NP.

Proof

Guess the map $f : V \rightarrow R^+$, append axioms $\{ v \equiv f(v) \mid v \in V \}$ to $\Gamma$ to get $\Gamma^+$, and verify using the word problem algorithm that $x \equiv_{\Gamma^+} y$.

Theorem 18

Boolean satisfiability $\preceq^m_{\log}$ SATIS.

Proof

The proof is left to the reader.

Corollary 19

SATIS is $\preceq^m_{\log}$-complete for NP.

We have immediately that SATIS is complete for co-NP.

Unfortunately, this does not help us with the problem VALID.

Let $\mathcal{A}$ be presented by $\Gamma$.

Definition

Let $d \in N$. Define $w \downarrow_d z$ iff
\( \exists w^* \text{ such that } w^* \preceq z \) and the length of the path from the root of \( w^* \) to the root of \( z \) is at most \( d \), as illustrated below.

\begin{center}
\includegraphics[width=0.3\textwidth]{diagram.png}
\end{center}

Lemma 20

If \( \mathcal{A} \) is infinite, then for any \( d,k \), there exists a set of \( k \) arbitrarily large terms \( \{w_1, \ldots, w_k\} \) such that if \( i \neq j \) then \( w_i \) and \( w_j \) are \( \Delta_d \)-incomparable.

Proof

If \( k = 1 \), then since \( \mathcal{A} \) is infinite, an arbitrarily large \( w_1 \) can be chosen not congruent to any smaller term. Now suppose \( \{w_1, \ldots, w_{k-1}\} \) have been constructed so that

\( w_i \not\preceq_d w_j \) for \( i \neq j \). Take \( w_k^0 \) to be larger than all terms in \( R_1 \) and larger than all of \( w_1, \ldots, w_{k-1} \) such that \( w_k^0 \) is not congruent to any smaller term. Let \( \Theta \) be an operator, and define

\( w_{k+1} = \Theta w_k^m \ldots w_k^m \). Then \( \forall m, n \forall i < k \ w_k^m \not\preceq_n w_i \); if not, then

for some \( m \), \( w_k^0 \not\preceq w_k^m = w^* \preceq w_i \), and an analysis of a proof

\( w_k^m \not\preceq w^* \) shows that \( w_k^0 \) would necessarily be congruent to a smaller term. Moreover, \( \forall m \forall i < k, w_i \not\preceq_m w_k^m \), by induction on \( m \):

certainly \( w_i \not\preceq w_k^0 \) implies \( w_i \not\preceq w_k^0 \), contrary to our choice of
and if \( w_i \leq_m w_k \) then \( w_i \leq_{m+1} w_k \) implies \( w_i \equiv w_k \), which by analysis of a proof \( w_i \leq_{m+1} w_k \) can be shown impossible, as above. Now let \( w_k = w_i \). Then \( \{w_1, \ldots, w_k\} \) is desired set. 

**Theorem 21**

\( \text{VALID} \in \text{co-NP} \).

**Proof**

Let \( \Gamma \) present \( \mathcal{A} \) and let \( x \equiv y \) be a schema with variables \( V \). Use the algorithm of Theorem 12 to determine whether \( \mathcal{A} \) is finite. If so then \( A = \Gamma \), so verify in parallel for each assignment \( I : V \rightarrow R \) that \( x[V \setminus I(V)] \equiv y[V \setminus I(V)] \), using the word problem algorithm. If \( \mathcal{A} \) is infinite, let \( \Gamma^+ = \Gamma \) but consider \( \Gamma^+ \) as inducing a congruence over \( \tau^+ \) instead of \( \tau \), where \( \tau^+ = \{\text{terms over } G \cup V, 0\} \). Let \( R^+ = R_{\Gamma^+} \).

**Claim**

If \( \mathcal{A} \) is infinite, \( x \equiv y \) is valid iff \( x \equiv_{\Gamma^+} y \).

The claim establishes the result, since if \( \mathcal{A} \) is infinite we can use the word problem algorithm to check \( x \equiv_{\Gamma^+} y \), and the parallel procedure given above for the finite case is a \( \Pi^1_1 \) computation\(^6\), thus in co-NP.

**Proof of claim**

\( (+) \) is clear. To show \( (\rightarrow) \), suppose \( x \equiv y \) is valid. Let \( d > \max \{\text{height}(z) \mid z \in R^+\} \) and let \( V = \{v_1, \ldots, v_k\} \). The previous lemma guarantees arbitrarily large \( w_1, \ldots, w_k \) such that if \( i \neq j \) then \( w_i \not\equiv_d w_j \). Take such \( w_1, \ldots, w_k \) to
be each of height > d and let I(v_i) = w_i, kisk. Let

\[ x' = x[V \setminus I[V]] , \]
\[ y' = y[V \setminus I[V]] . \]

Then \( x' + y' \), but no ancestor of any \( w_i \) is ever the root of a transformation, thus \( x' + y' \) via a proof in which all transformations to terms \( \not\equiv \)-incomparable to any \( w_i \) are done first, followed by transformations to the interiors of the \( w_i \). Thus \( \exists z \ z' = z[V \setminus I[V]] , x \equiv^+_I z , \) and \( z' + y' \) via transformations only to the interiors of the \( w_i \). But since the \( w_i \) are so big, each \( w_i \) in \( z' \) must go to either a subterm or superterm of some \( w_j \) in \( y' \) in the proof \( z' + y' \). But this says that either \( w_i \not\equiv^+_d \) or \( w_j \not\equiv^+_d \) \( w_i \), hence it must be that \( i = j \). But then \( z' = y' \). Since for \( i \neq j \) we have \( w_i \not\equiv^+_d \) \( w_j \), all occurrences of \( w_i \) in \( z' \) or \( y' \) replaced an occurrence of \( v_i \) in \( z' \) or \( y' \), thus \( z = y \). Therefore \( x \equiv^+_I y \), and the claim is verified.

Theorem 22

VALID is \( \leq^m_{\log} \) hard for co-NP.

Proof

The reduction from the Boolean tautology problem is left to the reader.

Corollary 23

VALID is \( \leq^m_{\log} \) complete for co-NP.

It should be noted that if we allow quantification over variables, deciding membership in SATIS is equivalent to
deciding truth of closed formulas (those in which all variables are quantified) of the form

\[ \exists v_1 \ldots \exists v_n \exists y, \]

and deciding membership in \( \text{VALID} \) is equivalent to deciding

truth of closed formulas of the form

\[ \forall v_1 \ldots \forall v_n \exists x \exists y, \]

in the algebra presented by \( \Gamma \) (equivalently, in all algebras satisfying the axioms of \( \Gamma \)). This is quite remarkable in view of the fact that the quantified variables range unboundedly over a possibly infinite set. In other results of this type,\(^5,7\) either the structure is finite, or the quantifiers are bounded.

If we define \( S_n, V_n \) by

\[ S_n(V_n) = \{ \langle \Gamma, \bar{Q} \, x \equiv y \rangle \mid \bar{Q} \, x \equiv y \text{ is a closed formula where } \bar{Q} \text{ is a string of quantifiers with } n \text{ alternations, the outermost a } \exists(V), \bar{Q} \, x \equiv y \text{ is true in } ^{\Gamma} / _{\exists P} \} \]

and if we let \( \Sigma^n_P \) and \( \Pi^n_P \) represent the \( n^{\text{th}} \) \( \Sigma \) and \( \Pi \) levels of the polynomial time hierarchy,\(^7\) we have by the preceding results

i) \( S_0 = V_0 \) is \( z^m_{\log} \)-complete for \( \Sigma^0_P = \Pi^0_P = \mathbf{P} \);

ii) \( S_1 \) is \( z^m_{\log} \)-complete for \( \Sigma^1_P = \mathbf{NP} \);

iii) \( V_1 \) is \( z^m_{\log} \)-complete for \( \Pi^1_P = \text{co-NP} \).

It is conjectured that, like other results in this area,\(^6,7\)

\( S_n \) is \( z^m_{\log} \)-complete for \( \Sigma^n_P \) and \( V \) is \( z^m_{\log} \)-complete for \( \Pi^n_P \),

for all \( n \); and \( \bigcup_n S_n \cup V_n \) is complete for \( \mathbf{PSPACE} \).
5. Isomorphism of finitely presented algebras

In this section we wish to show that the problem of isomorphism of finitely presented algebras (ISOM) is polynomially equivalent to the problem of graph isomorphism.

As before, the reduction from the graph problem (the more specific) to the algebra problem (the more general) is trivial. To go the other way, we show that every finitely presented algebra has a "reduced" presentation, which is unique in a certain well defined sense. In view of the relationship between finitely presented algebras and regular tree languages noted in §1, this result corresponds roughly to the minimization of states in a finite tree automaton.

In proving ISOM $\subseteq^m_P$ graph isomorphism, we use the reduction sequence

isomorphism of undirected graphs without multiple edges or loops

$\subseteq^m_P$ isomorphism of labeled directed graphs

$\subseteq^m_P$ isomorphism of directed graphs

$\subseteq^m_P$ isomorphism of directed graphs without multiple edges or loops

$\subseteq^m_P$ isomorphism of undirected graphs without multiple edges or loops.
The $\leq^m_{\text{log}}$ reductions in the above sequence are easy exercises and are left to the reader.

Definition

\[
\text{ISOM} = \{<\Gamma, \Delta> \mid \Gamma \text{ and } \Delta \text{ present isomorphic algebras}\}.
\]

We will assume that the number of operator symbols of each arity in $O_\Gamma$ and $O_\Delta$ is the same. The interested reader may verify that there is a polynomial time algorithm to check whether two operator symbols in any $O$ specify the same operation, hence the assumption is without loss of generality.

To prove the reduction $\text{ISOM} \leq^m_P$ graph isomorphism, we will show that every finitely presented algebra $\mathcal{A}$ has a "reduced" presentation $\Gamma$, which can be found in polynomial time, such that $r_\Gamma$ is unique (up to isomorphism). But $r_\Gamma$ is uniquely represented by the characteristic graph $\chi_\Gamma$ introduced in §3, as shown by the following lemma:

Lemma 24

$r_\Gamma$ and $r_\Delta$ are isomorphic (as subsets of algebras)

iff $\chi_\Gamma$ and $\chi_\Delta$ are isomorphic (as graphs).

Proof

Follows directly from Lemma 10.

Definition

A presentation is reduced provided
i) if \( \theta x_1 \ldots x_m \equiv \theta y_1 \ldots y_m \) is an axiom, then one of
   \( x_1 \not\in y_1, \ldots, x_m \not\in y_m \).

ii) no axiom of the form \( a \equiv x \) occurs, where \( a \in G \) and
    \( a \) does not occur in \( x \).

\( \square \)

**Lemma 25**

There is a polynomial time algorithm which for input \( \Gamma \) gives an equivalent reduced \( \Gamma^* \).

**Proof**

Given \( R_\Gamma \), find all congruent pairs, using the word problem algorithm of Theorem 1. Repeat the following two steps until no more changes occur:

a) If \( \theta x_1 \ldots x_m \equiv \theta y_1 \ldots y_m \) is an axiom and
   \( x_1 \equiv y_1, \ldots, x_m \equiv y_m \) follow, replace
   \( \theta x_1 \ldots x_m \equiv \theta y_1 \ldots y_m \) in \( \Gamma \) with new axioms
   \( x_1 \equiv y_1, \ldots, x_m \equiv y_m \).

b) If \( a \equiv x \) is an axiom, \( a \in G \), and \( a \) does not occur in \( x \), replace all occurrences of \( a \) in other terms of \( R_\Gamma \) with pointers to \( x \), and eliminate the axiom \( a \equiv x \).

\( \text{E.g.} \)

```
1 2
a a  \equiv \triangle x
```

becomes

```
1 2
\triangle x
```

```
1 2
\triangle x
```
We claim first that this algorithm is polynomially time bounded. Note that step b) occurs at most \( n \) times, since each time, a generator symbol disappears. For each occurrence of b), step a) occurs at most \( n^2 \) times, provided whenever
\[
\theta x_1 \ldots x_m \equiv \theta y_1 \ldots y_m \text{ and } \theta z_1 \ldots z_k \equiv \theta w_1 \ldots w_k \text{ are axioms, and}
\]
\[
x_1 \equiv y_1', z_j \equiv w_j \text{ for } i \neq m, j \neq k, \text{ and } \theta z_1 \ldots z_k < \theta x_1 \ldots x_m,
\]
step a) is applied to the axiom \( \theta x_1 \ldots x_m \equiv \theta y_1 \ldots y_m \) first. We must also insure that every time b) occurs, a valid presentation results, i.e. the graph remains acyclic. This follows from the requirement in b) that a not occur in \( x \) in the axiom \( a \equiv x \).

Hence the algorithm halts in polynomial time with a reduced presentation \( \Gamma^* \), so it remains to show that \( \Gamma^* \) and \( \Gamma \) are equivalent, i.e. present the same algebra. If a) is applied, we have axiom \( \theta x_1 \ldots x_m \equiv \theta y_1 \ldots y_m \in \Gamma \) and
\[
x_1 \equiv y_1', \ldots, x_m \equiv y_m. \text{ Let } \Gamma' = \Gamma - \{ \theta x_1 \ldots x_m \equiv \theta y_1 \ldots y_m \}. \]
Since \( \theta x_1 \ldots x_m \equiv \theta y_1 \ldots y_m \) follows from \( \Gamma' \cup \{ x_1 \equiv y_1, \ldots, x_m \equiv y_m \} \), we have that under the assumptions \( \Gamma' \), \{ \( x_1 \equiv y_1, \ldots, x_m \equiv y_m \} \) and \( \{ \theta x_1 \ldots x_m \equiv \theta y_1 \ldots y_m \} \) are equivalent. If b) is applied, let \( a \equiv x \) be the axiom removed. Let \( \tau' = \{ \text{terms over } O \& G(a) \} \), and let \( f: \tau + \tau' \) be given by \( f(y) = y[a \setminus x] \). An application of b) replaces axioms \( \Gamma = \{ x_1 \equiv y_1, \ldots, x_k \equiv y_k \} \) with
\[
\Gamma' = \{ f(x_1) \equiv f(y_1), \ldots, f(x_k) \equiv f(y_k) \}, \text{ inducing congruence}
\]
\( \equiv' = \equiv_{\Gamma'} \), so we need to show that \( \tau/\equiv \) and \( \tau'/\equiv' \) are isomorphic. But it is easily verified that for \( z, y \in \tau' \), \( z \equiv y \iff z \equiv y \), and each \( y \in \tau - \tau' \) is congruent via \( \equiv \) to \( y[a \setminus x] \in \tau' \), thus there exists an \( h \) such that the diagram
commutes, and \( h \) is an isomorphism.

In the following, let \( \Gamma \) and \( \Delta \) be finite presentations of algebras \( A \) and \( B \), respectively. The symbols \( G, 0, \tau \), etc. will have their usual meaning, except we will attach subscripts \( \Gamma \) and \( \Delta \) to denote the presentation with which they are associated.

**Lemma 26**

Suppose \( A \) and \( B \) are isomorphic via \( h \), and suppose \( \Gamma \) is reduced. Then there is a function \( f: \tau_\Gamma \to \tau_\Delta \) such that

i) the diagram

\[
\begin{array}{ccc}
\tau_\Gamma & \xrightarrow{f} & \tau_\Delta \\
\downarrow \scriptstyle{[\.]}_\Gamma & & \downarrow \scriptstyle{[\.]}_\Delta \\
\tau_\Gamma/\Xi_\Gamma & \xrightarrow{h} & \tau_\Delta/\Xi_\Delta
\end{array}
\]

commutes, and

ii) \( f[R_\Gamma] \subseteq R_\Delta \).

**Proof**

We have assumed previously that the number of operators of each arity in \( O_\Gamma \) and \( O_\Delta \) are the same. Since \( A \) and \( B \) are
isomorphic, there is a 1-1 arity preserving correspondence between $O_\Gamma$ and $O_\Delta$. Hence for notational convenience and without loss of generality, we will assume $O_\Gamma = O_\Delta$ and the correspondence is given by identity.

We first define an $f_1 : \tau_\Gamma \rightarrow \tau_\Delta$ satisfying i) only. For $a \in G_\Gamma$, let $f_1(a) = y$ where $y$ is any term in $\tau_\Delta$ such that $[y]_\Delta = h([a]_\Gamma)$. $f_1$ extends uniquely to a homomorphism $\tau_\Gamma \rightarrow \tau_\Delta$, by taking $f_1(\theta x_1 ... x_m) = \theta f_1(x_1) ... f_1(x_m)$. Then $f_1$ satisfies i), since for $a \in G$, $[f_1(a)]_\Delta = h([a]_\Gamma)$ by definition, and proceeding by structural induction,

\[
[f_1(\theta x_1 ... x_m)]_\Delta = \theta [f_1(x_1)]_\Delta ... [f_1(x_m)]_\Delta
= \theta h([x_1]_\Gamma) ... h([x_m]_\Gamma)
= h(\theta [x_1 ... x_m]_\Gamma),
\]

since $[,]_\Delta \circ f_1$ and $h \circ [,]_\Gamma$ are homomorphisms.

Let us take $f(x) = f_1(x)$ for $x \notin R_\Gamma$. We need only to find, for every $x \in R_\Gamma$, a $y \in R_\Delta$ such that $y \equiv_\Delta f_1(x)$; then we can take $f(x) = y$, and ii) will hold, but $f$ will satisfy i), since $\forall x f(x) \equiv_\Delta f_1(x) \rightarrow \forall x [f(x)]_\Delta = [f_1(x)]_\Delta = h([x]_\Gamma)$.

**Case 1:** $x$ appears in an axiom $x \equiv y \in \Gamma$, and neither $x$ nor $y$ are generator symbols.

Observe that since $f_1$ satisfies i), $x \equiv_\Gamma y \leftrightarrow f_1(x) \equiv_\Delta f_1(y)$.

Let $f_1(x) \star f_1(y)$ be a proof of $f_1(x) \equiv_\Delta f_1(y)$. If the entire tree is ever replaced, then we have $f_1(x) \star z \star w \star f_1(y)$ where $z \equiv w$ is an axiom of $\Delta$. Then $z \in R_\Delta$ hence we can take $f(x) = z$, and we are done. Otherwise, since neither $x$ nor $y$ are generator symbols, by construction of $f_1$ we have
\( f_1(x) = \theta f_1(x_1) \ldots f_1(x_m), \ f_1(y) = \theta f_1(y_1) \ldots f_1(y_m), \) and
\( f_1(x_1) \cong f_1(y_1), \ldots, \ f_1(x_m) \cong f_1(y_m), \) where \( x = \theta x_1 \ldots x_m \)
and \( y = \theta y_1 \ldots y_m. \) But then \( x_1 \cong y_1, \ldots, \ x_m \cong y_m, \) contradicting the fact that \( \Gamma \) was reduced.

**Case 2:** \( a \in G_\Gamma \) and \( a \) occurs in an axiom \( a \cong x, \) where \( x \not\in G_\Gamma. \)

Since \( \Gamma \) is reduced, \( x \) must contain occurrences of \( a. \) Then \( f_1(a) \cong f_1(x) \) and \( f_1(a) \neq f_1(x). \) Let \( y \) be a \(-\)minimal element of \( \{f_1(a)\}_\Delta, \) and let \( w = f_1(x)[f_1(a) \setminus y]. \) Then \( w \vdash y \) and \( y \nvdash w. \) In a proof \( w \vdash y, \) either an ancestor of some occurrence of \( y \) in \( w \) is the root of a transformation, in which case \( y \cong \Delta u \gamma R_\Delta \) as above; or not, in which case \( y \) is congruent to a proper subterm of itself, contradicting the assumption of \(-\)minimality.

**Case 3:** \( x \) is a proper subterm of a term \( y \) appearing in an axiom \( y \cong z. \)

By cases 1 and 2, we have \( f(y) \in R_\Delta. \) But then \( f_1(y) \cong f(y) \) and \( f_1(x) \neq f_1(y). \) In a proof \( f_1(y) \vdash f(y), \) either an ancestor of \( f_1(x) \) is the root of a transformation or not. If so, \( f_1(x) \) is congruent to a subterm of an axiom of \( \Delta, \) if not, \( f_1(x) \) is congruent to a subterm of \( f(y). \) But in either case, \( \exists w \in R_\Delta \)
\( f_1(x) \cong w, \) hence we can take \( f(x) = w. \)
Case 4: \( a \in G_{\Gamma} \) and \( a \) occurs only in the axiom \( a \equiv a \). Then
\[
 f_1(a) \text{ must be in } G_{\Delta}, \text{ otherwise } \exists x_1 \ldots x_n \in \tau_{\Gamma} \text{ such that }
 f_1(a) \equiv_{\Delta} \theta f_1(x_1) \ldots f_1(x_n), \text{ since } h \text{ is an isomorphism; but }
\]
then \( a \equiv_{\Gamma} \theta x_1 \ldots x_n \), which is impossible if \( a \) occurs only in \( a \equiv a \). Thus take \( f(a) = f_1(a) \in R_{\Delta} \).

Theorem 27

Let \( \mathcal{A}, \mathcal{B} \) be isomorphic via \( h \), and let \( \Gamma \) and \( \Delta \) be reduced.

Then \( r_{\Gamma} \) and \( r_{\Delta} \) are isomorphic via \( h \).

Proof

Using the lemma, form \( f \) and \( g \) such that

\[
\begin{array}{c}
\tau_{\Gamma} \\
\downarrow \quad f \\
[\ ]_{\Gamma} \\
\downarrow h \\
\mathcal{A} \\
\end{array}
\quad \quad
\begin{array}{c}
\tau_{\Delta} \\
\downarrow g \\
[\ ]_{\Delta} \\
\downarrow h^{-1} \\
\mathcal{B} \\
\end{array}
\quad \quad
\begin{array}{c}
\tau_{\Gamma} \\
\downarrow \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \quad \qua
Proof

Define $f : \tau_{\Gamma} \to \tau_{\Delta}$ by taking $f(a) = y$ where $y \in R_{\Delta}$ such that $[y]_{\Delta} = h([a]_{\Gamma})$ for $a \in G_{\Gamma}$, and extend $f$ to the unique homomorphism $\tau_{\Gamma} \to \tau_{\Delta}$, as in the previous proof.

Claim

\[ \begin{array}{c}
R_{\Gamma} \\
\downarrow h \\
\tau_{\Gamma}
\end{array} \quad \begin{array}{c}
f[R_{\Gamma}] \\
\downarrow \\
[ ]_{\Delta}
\end{array} \quad \begin{array}{c}
\tau_{\Delta}
\end{array} \]

commutes.

Proof of claim

For $a \in G_{\Gamma}$, $h([a]_{\Gamma}) = [f(a)]_{\Delta}$ by definition, and for $\emptyset x_1 \ldots x_m \in R_{\Gamma}$, we have $x_1, \ldots, x_m \in R_{\Gamma}$, hence by structural induction,

\[
h([\emptyset x_1 \ldots x_m]_{\Gamma}) = \emptyset h([x_1]_{\Gamma}) \ldots h([x_m]_{\Gamma})
\]

\[
= \emptyset [f(x_1)]_{\Delta} \ldots [f(x_m)]_{\Delta}
\]

\[
= [f(\emptyset x_1 \ldots x_m)]_{\Delta},
\]

and the claim is verified.

We wish to extend $h$ to $\hat{h}$ on domain $\mathcal{A}$ by taking

\[
\hat{h}([x]_{\Gamma}) = [f(x)]_{\Delta},
\]

but first we must show that $[ ]_{\Delta} \circ f$ is well-defined on $\Xi_{\Gamma}$-congruence classes, i.e. $x \Xi_{\Gamma} y \Rightarrow [f(x)]_{\Delta} = [f(y)]_{\Delta}$, so that $\hat{h}$ will be well-defined.

For $x, y \in \tau_{\Gamma}$, take $x \bar{=} y$ iff $[f(x)]_{\Delta} = [f(y)]_{\Delta}$.

Since $[ ]_{\Delta} \circ f$ is a homomorphism, $\bar{=}$ is a congruence relation on $\tau_{\Gamma}$. By the above claim, we have for $x, y \in R_{\Gamma}$
\[ x \equiv \Gamma y \iff [x]_{\Gamma} = [y]_{\Gamma} \]
\[ \text{iff } h([x]_{\Gamma}) = h([y]_{\Gamma}) \]
\[ \text{iff } [f(x)]_{\Delta} = [f(y)]_{\Delta} \]
\[ \text{iff } x \equiv \vec{z} y. \]

Since \( \equiv \Gamma \) is defined to be the smallest (most general) congruence satisfying the axioms of \( \Gamma \), and \( \vec{z} \) satisfies the axioms of \( \Gamma \), it follows that \( \equiv \Gamma \) is a refinement of \( \vec{z} \). Hence \( \forall x, y \in \equiv \Gamma \),
\[ x \equiv \Gamma y \Rightarrow x \vec{z} y + [f(x)]_{\Delta} = [f(y)]_{\Delta}, \]

as was to be shown.

Now we have that

\[
\begin{array}{c}
\Delta \\
\downarrow \quad \hat{h} \\
\Delta \\
\end{array}
\]

commutes, and \( \hat{h} \Vdash \Gamma = h \). We can also form \( g \) such that

\[
\begin{array}{c}
\Gamma \\
\downarrow \quad \hat{h}^{-1} \\
\Gamma \\
\end{array}
\]

commutes, and \( \hat{h}^{-1} \Vdash \Delta = h^{-1} \). But now \( \forall x \in \Gamma \),
\[ g \circ f(x) \equiv \Gamma x, \text{ by structural induction: certainly for } a \in G_{\Gamma}, \]
\[ [g \circ f(a)]_{\Gamma} = \hat{h}^{-1}([f(a)]_{\Delta}) \]
\[ = h^{-1}(h([a]_{\Gamma})) \]
\[ = h^{-1}(h([a]_{\Gamma})) \]
\[ = [a]_{\Gamma}, \text{ and then} \]
\[ [g^*f(0x_1 \ldots x_m)] \Gamma = 0[g^*f(x_1)] \Gamma \ldots [g^*f(x_m)] \Gamma \]
\[ = 0[x_1] \Gamma \ldots [x_m] \Gamma \]
\[ = [0x_1 \ldots x_m] \Gamma. \]

Similarly \( \forall x \in \Gamma \), \( f^*g(x) \equiv \Delta \). But this says that \( \hat{h} \) and \( \hat{h}^{-1} \) are inverses, since \( \forall x \in \Gamma \)
\[ \hat{h}^{-1}(\hat{h}([x] \Gamma)) = \hat{h}^{-1}(f([x]) \Delta) \]
\[ = [g^*f(x)] \Gamma \]
\[ = [x] \Gamma \]
and similarly \( \forall y \in \Delta \), \( \hat{h}(\hat{h}^{-1}([y] \Delta)) = [y] \Delta \). Thus \( \mathcal{A} \) and \( \mathcal{B} \)
are isomorphic via \( \hat{h} \).

\textbf{Corollary 29}

\( ISOM_{P}^{m} \) isomorphism of labeled directed graphs.

\textbf{Proof}

Given an instance of \( ISOM_{< \Gamma, \Delta> } \), reduce \( \Gamma \) and \( \Delta \) to get \( \Gamma^{*} \) and \( \Delta^{*} \), and then form the graphs \( \chi_{\Gamma}^{*} \) and \( \chi_{\Delta}^{*} \). By Lemmas 10 and 25, this can be done in polynomial time. Then by Theorems 27 and 28 and Lemma 24,

\[ \mathcal{A} \cong \mathcal{B} \text{ iff } r_{\Gamma}^{*} = r_{\Delta}^{*} \text{ iff } \chi_{\Gamma}^{*} = \chi_{\Delta}^{*}. \]

Again, it is rather remarkable that isomorphism of possibly infinite structures should reduce to that of finite ones, unlike previous results in this area.15,16

The \( \equiv_{P}^{m} \) equivalence of graph isomorphism and \( ISOM \) should be of great interest to those who believe graph isomorphism is \( NP \) complete. It is clear that graph isomorphism is in \( NP \), but the form of the problem is so restricted (i.e., two graphs
with the same numbers of vertices of each in- and out-degree) that standard reduction techniques fail to show even that it is hard for P. However, there are no such restrictions on the form of instances of ISOM, and it is quite trivial to show ISOM is P-hard:

**Theorem 30**

ISOM is \( \sqrt{\log} \)-hard for P.

**Proof**

Use \( \Gamma' \) of Theorem 5 and a presentation of the trivial algebra.

Thus it would surely be easier to show that ISOM, the more general of the two problems, is NP-hard.
6. Open questions

The following is a list of interesting and relevant open problems:

1) Prove the conjecture at the end of §4, thus generalizing the results of [7]. The techniques of §4 do not immediately generalize, nevertheless it is likely that this problem will be easier than 2) or 3).

2) Show graph isomorphism is $\leq_m^p$-complete for NP, by showing ISOM is $\leq_m^p$-hard for NP.

3) In lieu of 2), improve the $\leq_m^p$ reduction from ISOM to graph isomorphism to a $\leq_m^{\log}$ reduction, thereby showing graph isomorphism is $\leq_m^{\log}$-hard for P. The $\leq_m^p$ reduction given involves an algorithm for a problem complete for P, thus it will not be altered easily to a $\leq_m^{\log}$ reduction (unless LOGSPACE = P).

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References


