THE COMPLEXITY OF EVALUATION RELATIONAL QUERIES

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Abstract

We show that, given a relation $R$, a relational query $\varphi$ involving only projection and join, and a conjectured result $r$, testing whether $\varphi(R)=r$ is $D^P$-complete. Bounding the size of $\varphi(R)$ from below (above) is $NP$-hard (co-$NP$-hard), and bounding it both ways is $D^P$-hard. Computing the size of $\varphi(R)$ is $\#P$-hard.

We also show that, given two relations $R_1$ and $R_2$ and two queries $\varphi_1$ and $\varphi_2$ as above, testing whether $\varphi_1(R_1) \subseteq \varphi_2(R_2)$ and testing whether $\varphi_1(R_1)=\varphi_2(R_2)$ are both $\Pi^P_2$-complete, even when $R_1=R_2$ or when $\varphi_1=\varphi_2$.

Keywords: Relational database, relational expression, polynomial transformation, completeness, hardness, $NP$, co-$NP$, $D^P$, enumeration problem, polynomial-time hierarchy.

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1. Introduction

The relational model for databases [3] has been proposed as a formal means to describe data organization and to express queries. In this paper we are using the relational model to formalize the problem of evaluating a database query and the related problem of computing (or estimating) the size of the result, and we are using complexity theory to characterize the complexity of these problems. Specifically, we prove the following:

(i) Given a database \( s \), a relational query \( Q \) and a conjectured result \( r \), testing whether \( Q(s) = r \) is complete for the class \( D^P \) (the class of languages which are equal to the intersection of a language in \( NP \) and a language in co-\( NP \) --see [10, 6] for the relevant definitions). The problem is known to be \( NP \)-complete if we replace "=" by "\( \supseteq \)" [16], and co-\( NP \)-complete if we replace "=" by "\( \subseteq \)" [9]; we also give direct proofs of these facts.

(ii) Given a database \( s \), a query \( Q \) and two non-negative integers \( d_1, d_2 \), testing whether \( d_1 \leq |Q(s)| \leq d_2 \) (where \( |Q(s)| \) is the number of tuples of \( Q(s) \)) is \( D^P \)-hard, even when \( d_1 = d_2 \) or when \( d_1 < d_2 \). Testing whether \( d_1 \leq |Q(s)| \) is \( NP \)-hard, and testing whether \( |Q(s)| \leq d_2 \) is co-\( NP \)-hard (the co-\( NP \)-hardness result also follows from the co-\( NP \)-completeness result of [9]).

(iii) Given a database \( s \) and a query \( Q \), the enumeration problem of counting the number of tuples in the result \( Q(s) \) is \( \#P \)-hard (see [13, 6] for relevant definitions).

We also prove the following results, related to the behavior of relational queries when viewed as mappings (as in [2]):

(iv) Given a database \( s \) and two queries \( Q_1, Q_2 \), testing whether \( Q_1(s) \subseteq Q_2(s) \) and testing whether \( Q_1(s) = Q_2(s) \) are both complete for the class \( \Pi^P_2 \) of the polynomial-time hierarchy (see [11, 6] for relevant definitions).

(v) Given two databases \( s_1, s_2 \) and a query \( Q \), testing whether \( Q(s_1) \subseteq Q(s_2) \) and testing whether \( Q(s_1) = Q(s_2) \) are both \( \Pi^P_2 \)-complete.

In all the above results, databases can be constrained to consist of a single relation, and queries are restricted to only use the operations of projection and (natural) join.
2. Basic Definitions

The relational database model [3] assumes that the data are stored in tables called relations. The columns of a table correspond to attributes, and the rows to tuples (records). Each attribute $A$ has an associated domain of values $\text{Dom}(A)$. A relation scheme is a finite set of attributes labeling the columns of a table, and it is usually written as a string of attributes. Let $X$ be a relation scheme: an $X$-tuple is a mapping $\mu$ from $X$ into $\bigcup A \in X \text{Dom}(A)$ such that $\mu(A) \in \text{Dom}(A)$ for each attribute $A$ in $X$; a relation over $X$ is a finite set of $X$-tuples. A database scheme $S$ is a finite set of relation schemes. A database over $S$ is a set of relations containing exactly one relation over each relation scheme in $S$.

In the context of the relational model, one way of formulating queries is by using a set of operations defined on relations (relational algebra [3, 4]). In this paper we only consider two operations, projection and join.

The projection $[Y]$ of an $X$-tuple $t$ onto a subset $Y$ of $X$ is the restriction of $t$ to $Y$. The projection $\pi_Y(R)$ of a relation $R$ over $X$ to $Y$ is the set of projections of the tuples in $R$ to $Y$. If $R_1$, $R_2$ are relations over the relation schemes $X_1$ and $X_2$ respectively, the join of $R_1$ and $R_2$, written $R_1 \Join R_2$, is the relation $R_1 \Join R_2 = \{t \mid t \in X_1 \cup X_2\text{-tuple, } \mu[X_1] \in R_1, \mu[X_2] \in R_2\}$. The join of a finite set of relations $\{R_i\}$ will be written as $\Join R_i$.

A relational expression consists of relation schemes as operands and projection and join as operations. A relational expression $\varphi$ defines, in the obvious way, a function which takes one argument for each relation scheme $X$ appearing in the expression as operand (the corresponding argument is a relation over $X$), and produces as result a relation over a certain relation scheme, the target relation scheme of $\varphi$, $\text{tr}x(\varphi)$. We will be using relational expressions to formulate queries in relational databases.

We also give the definition of the complexity class $\mathcal{D}^P$: $\mathcal{D}^P = \{L_1 \cap L_2 : L_1 \in \mathcal{NP}, L_2 \in \mathcal{co-NP}\}$. We have $\mathcal{NP} \cup \mathcal{co-NP} \subseteq \mathcal{D}^P \subseteq P^{\mathcal{NP}}$ (polynomial time with an oracle from $\mathcal{NP}$). We also note that, unless $\mathcal{NP} = \mathcal{co-NP}$, a problem which is complete for $\mathcal{D}^P$ is not in $\mathcal{NP} \cup \mathcal{co-NP}$. For a discussion of the type of languages of which $\mathcal{D}^P$ is the natural niche, and also for more natural problems
which turn out to be complete for $D^P$, see [10].

3. $D^P$-completeness and $\#P$-hardness Results

We will first prove the following:

**Theorem 1:**

Given a relation $R$, a relational expression $\varphi$ over projection and join, and a relation $r$, it is $D^P$-complete to test whether $\varphi(R) = r$.

**Theorem 2:**

Given a relation $R$, a relational expression $\varphi$ over projection and join, and two "small" non-negative integers $d_1, d_2$ (i.e. written in unary), it is $D^P$-complete to test whether $d_1 \leq |\varphi(R)| \leq d_2$, even when $d_1 = d_2$ or when $d_1 < d_2$. Testing whether $d_1 \leq |\varphi(R)|$ is $NP$-complete, and testing whether $|\varphi(R)| \leq d_2$ is $co-NP$-complete.

All the reductions in this paper are reminiscent of the reduction in [1]. The reductions used to prove the $D^P$-completeness results in Theorem 1 and Theorem 2 are from the following problem:

**3SAT-3UNSAT:** "Given two Boolean expressions $G, G'$ in 3-conjunctive normal form (3CNF), is it true that $G$ is satisfiable and $G'$ is not?"

It is immediate that 3SAT-3UNSAT is in $D^P$, and it is also a straightforward consequence of the fact that the 3-satisfiability problem is $NP$-complete [5, 6, 8] that 3SAT-3UNSAT is complete for $D^P$ (see also [10]). Furthermore, it is clear that we may restrict ourselves to expressions containing at least three clauses, and such that the variables appearing in each clause are distinct.

Now let $G = F_1 \ldots F_m$ be a Boolean expression in 3-conjunctive normal form; the $F_j$'s are clauses of three literals each and the variables appearing in the expression are $x_1, x_2, \ldots, x_n$. We denote the variables appearing in a clause $F_j$ by $x_{j_1}, x_{j_2}, x_{j_3}$. We construct a relation $R_G$ corresponding to $G$ as follows: $R_G$ has $n+1+m(m+1)/2$ columns. The first $m$ columns correspond to the clauses of $G$ and are labeled by the attributes $F_1, \ldots, F_m$; the next $n$ columns correspond to the variables in $G$ and
are labeled by the attributes $X_1, X_2, \ldots, X_n$; the next $m(m-1)/2$ columns correspond to the two-element subsets of $\{1, \ldots, m\}$, and are labeled by the attributes $Y_{\{1,2\}}, \ldots, Y_{\{1,m\}}, \ldots, Y_{\{m-1,m\}}$; the last column is labeled by the attribute $S$.

For each clause $F_j$ of $G$, $R_G$ has 7 tuples as follows: let $h_{jk}$, $k=1, \ldots, 7$, be the seven satisfying truth assignments of the clause $F_j$ (each $h_{jk}$ is a function from $\{x_{j1}, x_{j2}, x_{j3}\}$ to $\{0, 1\}$); for each $h_{jk}$, $R_G$ contains a tuple $\mu_{jk}$ such that $\mu_{jk}(x_{ji})=1$, $\mu_{jk}(x_{ji})=e$ for $i \neq j$, $\mu_{jk}(X_i)=h_{jk}(x_{j1})$, $i=1, 2, 3$, $\mu_{jk}(X_i)=e$ for $i \neq j$, $i=1, 2, 3$, $\mu_{jk}(Y_{\{i,j\}})=x$ if $j=i$ or $j=l$ and $\mu_{jk}(Y_{\{i,j\}})=e$ otherwise, $\mu_{jk}(S)=a$. Finally, $R_G$ contains a tuple $\nu$, where $\nu(F_j)=1$, $\nu(S)=b$, $\nu(W)=e$ for $W \neq F_j$.

We also consider the following relational expression $\varphi_G$ corresponding to $G$:

$$\varphi_G = \pi_{F_1 \ldots F_m}(T) \ast [\pi_{F_1 \ldots F_m}(X_1 X_2 X_3 Y_{\{1,2\}} Y_{\{1,3\}} Y_{\{2,3\}} S(T))]$$

where $T$ is the relation scheme of $R_G$, namely $F_1 \ldots F_m X_1 X_2 X_3 Y_{\{1,2\}} \ldots Y_{\{1,m\}} \ldots Y_{\{m-1,m\}} S$.

**Example.** Let $G=(x_1+x_2+x_3)(\neg x_2+x_3+\neg x_4)(\neg x_3+\neg x_4+\neg x_5)$ ($\neg x$ stands for the negation of the variable $x$); the relation $R_G$ is

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The relational expression $\varphi_G$ is

$$\pi^* F_1 X_1 X_2 X_3 Y_{\{1,2\}} Y_{\{1,3\}} S^{*\pi} F_2 X_2 X_3 X_4 Y_{\{1,2\}} Y_{\{2,3\}} S^{*\pi} F_3 X_3 X_4 X_5 Y_{\{1,3\}} Y_{\{2,3\}} S$$

We remark at this point that the fact that the same symbols (0, 1, $e$, $x$) are used in different columns is irrelevant; one could imagine replacing (in a consistent way, of course) any symbols in any particular column by new symbols, appearing only in that column.

$R_G$ and $\varphi_G$ can be constructed in time polynomial in the space needed to write down $G$; they capture the satisfiability (or unsatisfiability) of $G$ as described below (let $F$ denote the relation scheme $F_1 \ldots F_m$ and $Y$ denote the relation scheme $Y_{\{1,2\}} \ldots Y_{\{1,m\}} \ldots Y_{\{m-1,m\}}$):

**Lemma 1:**

$\varphi_G(R_G) = R_G \cup R_{\varphi}$ where for each tuple $\mu$ in $R_{\varphi}$, $\mu(F_j) = 1$, $\mu(Y_{\{i,j\}}) = x$, $\mu(S) = a$, and $\mu[X_1 X_2 \ldots X_n]$ defines a satisfying truth assignment $h$ for $G$ (by taking $h(x_j) = \mu(X_j)$); conversely, for each satisfying truth assignment for $G$ there is a corresponding tuple in $R_{\varphi}$.

**Proof:** Let $\mu$ be a tuple in $\varphi_G(R_G)$. If $\mu(S) = b$, then $\mu = \nu$. If $\mu(S) = a$, then either $\mu[F] = \mu_{jk}[F]$ for some $j$, $k$, or $\mu[F] = \mu_F$. In the first case, it is not difficult to check that $\mu = \mu_{jk}$ (this is enforced by the $Y$ attributes); in the second case, it is clear that $\mu(F_j) = 1$, $\mu(Y_{\{i,j\}}) = x$, $\mu(S) = a$, and $\mu[X_1 X_2 \ldots X_n]$ goes through exactly the satisfying truth assignments for $G$.

**Proposition 1:**

If $G$ is unsatisfiable, $\pi \gamma \varphi_G(R_G) = \pi \gamma (R_G) \cup u_G$; if $G$ is satisfiable, $\pi \gamma \varphi_G(R_G) = \pi \gamma (R_G) \cup u_G$, where $u_G$ is the $Y$-tuple such that $u_G(Y_{\{i,j\}}) = x$.

**Proof:** Immediate, from Lemma 1.

The following is a simple fact which we will also be using in subsequent proofs:

**Proposition 2:**

Given a relation $R$, a relational expression $\varphi$ over projection and join, and a tuple $t$, testing whether $t \in \varphi(R)$ is in $NP$. 

Proof: Immediate, by an inductive argument on the structure of $\varphi$. Alternatively, one may consider the tableau [1] corresponding to $\varphi$, and guess a valuation showing that $t \in \varphi(R)$.

Notice that it immediately follows from Proposition 1 that $G$ is satisfiable iff $u_G \subseteq \pi \psi_G(R_G)$. Combining this fact with Proposition 2, we get a direct proof of the following result, proved indirectly in [16]:

"Given a relation $R$, a tuple $t$, and relation schemes $X$, $Y$, it is NP-complete to test whether $t \in \pi(X \ast \pi(Y(R)))$".

It also follows from Lemma 1 that $G$ is unsatisfiable iff $\varphi(G) = R_G$, and combining with Proposition 2 we get a direct proof of the following result [9]:

"Given a relation $R$ and relation schemes $Y$, testing whether $\ast \pi(Y(R)) = R$ is co-NP-complete".

We now prove Theorem 1 and Theorem 2.

Proof of Theorem 1:

We first show membership in $DP$: it suffices to show that testing whether $t \subseteq \varphi(R)$ is in NP, and testing whether $\varphi(R) \subseteq t$ is in co-NP. The first is a simple consequence of Proposition 2: just test whether $t \in \varphi(R)$ for all tuples $t \in R$. The second is equivalent to showing that testing whether $\varphi(R) \subseteq t$ is in NP ($\subseteq$ stands for the negation of $\subseteq$): for this, just nondeterministically guess a tuple $t$ and check that $t \in \varphi(R)$, and also check that $t \notin R$.

For the $DP$-hardness part, we will make a reduction from 3SAT-3UNSAT. Let $G, G'$ be two Boolean expressions in 3CNF. Let $R_G$ be the relation corresponding to $G$ over the relation scheme $T = F_1 \ldots F_m X_1 \ldots X_n Y_{\{1,2\}} \ldots Y_{\{1,m\}} \ldots Y_{\{m-1,m\}} S$, and $R_{G'}$ be the relation corresponding to $G'$ over the relation scheme $T' = F'_1 \ldots F'_{m'} X'_1 \ldots X'_{n'} Y'_{\{1,2\}} \ldots Y'_{\{1,m'\}} \ldots Y'_{\{m'-1,m'\}} S'$. Let $R_{G,G'} = R_G \ast R_{G'}$; observe that $T \cap T' = \emptyset$, and that for $Z \subseteq T$ we have $\pi(Z, R_{G,G'}) = \pi(Z, R_{G})$, and for $Z' \subseteq T'$ we have $\pi(Z', R_{G,G'}) = \pi(Z', R_{G'})$. Let $\varphi_{G,G'}$ be the relational expression $\pi(X \pi(Y(G \ast G'), taking as argument the relation scheme $T \cup T'$; clearly $\varphi_{G,G'}(R_{G,G'}) = \pi(Y \varphi(G(R_G)$
* \( \pi \gamma \varphi G(R_G') \), and thus using Proposition 1 it is easy to see that \( G \) is satisfiable and \( G' \) is unsatisfiable iff \( \varphi_{G,G'}(R_{G,G'}) = (\pi \gamma(R_G') \cup \varphi G) \ast \pi \gamma(R_G') \) (call this relation \( r_{G,G'} \)). Since \( R_{G,G'} \), \( \varphi_{G,G'} \), \( r_{G,G'} \) can be constructed in time polynomial in the space needed to write down \( G \), \( G' \), we are done.

Observe that in order to be able to make the reduction (and so in order for Theorem 1 to be true) it suffices to consider relational expressions of a certain restricted form. This will be the case for all of our results.

Proof of Theorem 2:

To prove the membership assertions, it suffices to show that testing whether \( d_1 \leq |\varphi(R)| \) is in \( NP \) and testing whether \( |\varphi(R)| \leq d_2 \) is in \( co\-NP \). The first follows from Proposition 2; just guess nondeterministically \( d_1 \) distinct tuples (recall that \( d_1 \) is in unary) and check that each of them is in \( \varphi(R) \). For the second it suffices to show that testing whether \( |\varphi(R)| \geq d_2 + 1 \) is in \( NP \), which is the same as the first.

For the \( D^P \)-hardness part let \( G, G' \) be as before, and let \( \beta = 7m + 1 \), \( \beta' = 7m' + 1 \). By appropriately padding \( G' \) (add a sufficient number of extra clauses not affecting its satisfiability) we can make sure that \( \beta < \beta' \). Now \( |\varphi_{G,G'}(R_{G,G'})| = |\pi \gamma \varphi G(R_G')| \cdot |\pi \gamma \varphi_{G,G'}(R_{G,G'})| \), and furthermore \( |\pi \gamma \varphi_{G,G'}(R_{G,G'})| = \beta \) if \( G \) is unsatisfiable and \( |\pi \gamma \varphi G(R_G')| = \beta + 1 \) if \( G \) is satisfiable, and similarly for \( G' \) (Proposition 1). Thus, \( G \) is satisfiable and \( G' \) is unsatisfiable iff \( |\varphi_{G,G'}(R_{G,G'})| = (\beta + 1) \beta' \), iff \( \beta (\beta' + 1) + 1 \leq |\varphi_{G,G'}(R_{G,G'})| \leq \beta (\beta' + 1) + \beta' \).

For the remaining hardness assertions it suffices to observe that, by Lemma 1, \( G \) is satisfiable iff \( \beta + 1 \leq |\varphi G(R_G')| \), and \( G \) is unsatisfiable iff \( |\varphi G(R_G')| \leq \beta \).

Using the constructions described, we can also show that counting the number of tuples in \( Q(s) \), where \( s \) is a database and \( Q \) a relational query, is at least as hard as counting the number of accepting computations of any polynomial time nondeterministic Turing machine. We make use of the fact that the problem of counting the number of satisfying truth assignments for an instance of 3-SATISFIABILITY is \( \# P \)-complete [14].
Theorem 3:

Given a relation $R$ and a relational expression $\varphi$ over projection and join, the enumeration problem of counting the number of tuples in $\varphi(R)$ is $\#P$-hard.

Proof: Immediate, by the fact that if $a(G)$ denotes the number of satisfying truth assignments for $G$, then $a(G) = |\varphi_G(R, G)| = 7m+1$ (Lemma 1).

By observing the form of $\varphi_G$ we can also see the following:

Corollary:

Given a relation $R$ and relation schemes $Y_i$, the enumeration problem of counting the number of tuples in $\pi_{Y_i}(R)$ is $\#P$-complete.

Proof: By the proof of Theorem 3, it suffices to show membership in $\#P$. Consider the following "counting Turing machine" $M$: $M$ nondeterministically guesses a tuple $t$, and accepts if and only if $[Y_i] \in R$ for all $i$. Clearly $M$ runs in polynomial time, and the number of accepting computations of $M$ is the same as the number of tuples in $\pi_{Y_i}(R)$.

4. $\Pi_2^P$-completeness Results

In this section we prove the following:

Theorem 4:

Given a relation $R$ and two relational expressions $\varphi_1$, $\varphi_2$ over projection and join, testing whether $\varphi_1(R) \subseteq \varphi_2(R)$ and testing whether $\varphi_1(R) = \varphi_2(R)$ are both $\Pi_2^P$-complete.

Theorem 5:

Given two relations $R_1$, $R_2$ and a relational expression $\varphi$ over projection and join, testing whether $\varphi(R_1) \subseteq \varphi(R_2)$ and testing whether $\varphi(R_1) = \varphi(R_2)$ are both $\Pi_2^P$-complete.

We first prove that the problems above are in $\Pi_2^P$. 
Proposition 3: Given two relations \( R_1, R_2 \) and two relational expressions \( \varphi_1, \varphi_2 \) over projection and join, testing whether \( \varphi_1(R_1) \subseteq \varphi_2(R_2) \) and testing whether \( \varphi_1(R_1) = \varphi_2(R_2) \) are both in \( \Pi^P \).

Proof: For the first part, it suffices to show that testing whether \( \varphi_1(R_1) \subseteq \varphi_2(R_2) \) is in \( \Sigma^P = NP^{NP} \) (nondeterministic polynomial time with an oracle from \( NP \)). But by Proposition 2 all we have to do is nondeterministically guess a tuple \( t \) and check, by asking an appropriate oracle from \( NP \), that \( t \in \varphi_1(R_1) \) and \( t \notin \varphi_2(R_2) \). The second part is similar.

The reductions used to prove Theorem 5 and Theorem 6 are from the following problem, which was shown to be \( \Pi^P \)-complete in \([11, 15]\):

Q-3SAT: "Given a Boolean expression \( G \) in 3-conjunctive normal form and a partition of the variables in \( G \) into two sets \( X = \{x_1, \ldots, x_r\}, X' = \{x_{r+1}, \ldots, x_n\} \), determine whether for all assignments of truth values to the variables in \( X \) \( G \) is satisfiable, i.e. determine whether \( \forall X \exists X'(G(X, X') = 1) \) (we write \( \forall X \) for \( \forall x_1 \ldots \forall x_r \) etc.)."

We will actually need to impose the following technical restrictions on the set \( X \):

Proposition 4:

Q-3SAT is \( \Pi^P \)-complete even if \( X \) is not contained in any \( V_j \) where \( V_j \) is the set of variables appearing in the clause \( F_j \) and also \( X \) does not contain any \( V_j \).

Proof: To enforce the first restriction, just add to \( G \) the clauses \((v_1 + v_2 + v_3) \), \((v_4 + v_5 + v_6)\), where the \( v_i \)'s are new variables, and replace \( X \) by \( X \cup \{v_1, v_4\} \). If \( X \) contains \( V_j \) then the problem is trivial, because there is an assignment to the variables in \( V_j \) making \( F_j \) false, and thus \( \forall X \exists X'(G(X, X') = 1) \) is false.

In the following reductions, we will assume that the restrictions described above are satisfied.

We will use \( X \) to denote both the set \( \{x_1, \ldots, x_r\} \) and the relation scheme \( X_1 \ldots X_r \).

Proof of Theorem 4:

By Proposition 3, we only need to prove the hardness part. Let \( G, X \) be as above, and let \( R'_G \) be a relation obtained from \( R_G \) as follows: first, for each clause \( F_j \) add a tuple \( \xi_j \) corresponding to
the truth assignment \( h_j: \{x_{j_1}, x_{j_2}, x_{j_3}\} \rightarrow \{0, 1\} \) which does not satisfy \( F_j \): that is, \( \xi(F_j) = 1 \), \( \xi(F_j) = e \) for \( i \neq j \), \( \xi(F_j) = h(x_j) \) for \( i = j \), \( i = 1, 2, 3 \), \( \xi(F_j) = e \) for \( i \neq j \), \( i = 1, 2, 3 \), \( \xi(F_j) = e \) for \( i = j \) or \( j = 1 \) and \( \xi(F_j) = e \) otherwise, \( \xi(F) = a \). Then add a new column labeled by \( U \), and make \( \xi(U) = e \) and \( \mu(U) = e \) for any other tuple \( u \) in the relation.

Consider also the following relational expressions (\( T \) is the relation scheme of \( R' \)) namely 

\[
F_1...F_mX_1X_2...X_nY_{\{1,2\}}...Y_{\{m-1,m\}}SU;
\]

\[
\varphi^1_G = \pi_{F_1...F_m(T)} \ast * \pi_{F_1X_1X_2...X_nY_{\{i\}}}S(T') \ast * \pi_U(T')
\]

\[
\varphi^2_G = \pi_{F_1...F_m(T)} \ast * \pi_{F_1X_1X_2...X_nY_{\{i\}}}S(T') \ast * \pi_{U(T')}
\]

It is not difficult to see that \( \varphi^2_G \) picks out the satisfying truth assignments for \( G \) (despite the extra tuples, since it looks at both the \( S \) and \( U \) columns), while \( \varphi^1_G \) considers \( G \) as a tautology (because of the extra tuples, since it only looks at the \( S \) column). More specifically, we have that

\[
\pi_X\varphi^1_G(R' G) = \pi_X(R' G) \cup R_X \quad \text{and} \quad \pi_X\varphi^2_G(R' G) = \pi_X(R' G) \cup R_{X,a}
\]

where:

a) \( R_X \) consists of all possible truth assignments to variables in \( X \), and \( R_{X,a} \) consists of the restrictions of the satisfying truth assignments for \( G \) to the variables in \( X \) (\( R_{X,a} = \pi_X(R_G) \)).

b) No tuple in \( \pi_X(R' G) \) may be taken as defining a truth assignment (by our first restriction on \( X \), any such tuple contains at least one \( e \)).

Thus, \( \forall X \exists X'(G(X, X') = 1) \) is true iff \( R_X \subseteq R_{X,a} \) iff \( \pi_X(R' G) \cup R_X \subseteq \pi_X(R' G) \cup R_{X,a} \) iff \( \pi_X\varphi^1_G(R' G) \subseteq \pi_X\varphi^2_G(R' G) \). This completes the proof.

Proof of Theorem 5:

By Proposition 3, we only have to prove the hardness part. Let \( G, X \) be as above, and let \( R'' G \) be a relation obtained from \( R' G \) by omitting the \( U \) column. We now have

\[
\pi_X\varphi_G(R'' G) = \pi_X(R'' G) \cup R_X \quad \text{and} \quad \pi_X\varphi_G(R G) = \pi_X(R G) \cup R_{X,a}
\]

As before, no tuple in either \( \pi_X(R'' G) \) or \( \pi_X(R G) \) may be taken as defining a truth assignment; moreover, we actually have \( \pi_X(R'' G) = \pi_X(R G) \) (by our second restriction on \( X \), the extra tuples in \( R'' G \) do not matter). Thus, \( \forall X \exists X'(G(X, X') = 1) \) is true iff \( R_X \subseteq R_{X,a} \) iff \( \pi_X(R'' G) \cup R_X \subseteq \pi_X(R G) \cup R_{X,a} \) iff \( \pi_X\varphi_G(R'' G) \subseteq \pi_X\varphi_G(R G) \), iff \( \pi_X\varphi_G(R'' G) = \pi_X\varphi_G(R G) \). This completes the proof.

Observe also that the following fact can be proved by a trivial modification to our proofs:
"Given a relation $R$, a relational expression $\varphi$ over projection and join, and for each attribute $X_i$ in $\text{trs}(\varphi)$, a finite subset $E_i$ of $\text{Dom}(X_i)$, it is $\Pi^P_2$-complete to test whether $\varphi(R)$ contains every possible tuple $\mu$ such that $\mu(X_i) \in E_i$ for all $X_i$ in $\text{trs}(\varphi)$".

5. Discussion

The results of Section 3 must be seen as further theoretical evidence supporting the intuitive fact that processing a query requires in general time exponential in the size of the query. Moreover, they imply that the intermediate results can be inherently much larger than both the input relation and the query output. They also provide strong evidence of the intractability of the problem of estimating the size of the query output.

The results of Section 4 characterize the complexity of testing the equivalence of relational queries on a fixed given relation (Theorem 4). They are also motivated by the theory of database mappings, which was proposed as a means to solve the view update problem [2]. If $S$ is a general state space of possible values of a database and $f$ is a mapping on $S$, one considers the partition of $S$ induced by the equivalence relation $\sim_f$ where $x \sim_f y$ iff $f(x) = f(y)$. Theorem 5 indicates the hardness of deciding this condition in the special case of relational mappings.

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References

(Note: References [7, 12] are not cited in the text.)


