Beyond Traktenbrot’s Theorem

- By Traktenbrot’s Theorem, it is undecidable to check whether a given first-order query $Q$ produces some output over some database.
- What happens if $D$ is actually given as input?

The following are natural (decision) problems in this context:

**QUERY-OUTPUT-TUPLE (QOT)**

**INSTANCE:** A database $D$, a query $Q$, a tuple $\vec{c}$ of values.

**QUESTION:** Does $\vec{c} \in Q(D)$ hold?

**BOOLEAN-QUERY-EVALUATION (BQE)**

**INSTANCE:** A database $D$, a Boolean query $Q$.

**QUESTION:** Does $Q$ evaluate to true in $D$?

**NOTE:** we often view Boolean domain calculus queries \{\emptyset|\phi\} simply as closed formulae $\phi$.

**QUERY-NON-EMPTINESS (QNE)**

**INSTANCE:** A database $D$, a query $Q$.

**QUESTION:** Does query $Q$ yield a non-empty result over the DB $D$, i.e. $Q(D) \neq \emptyset$?
QOT vs. BQE vs. QNE

We concentrate here mainly on the complexity of BQE.

Not a limitation: in our setting QOT and QNE are essentially the same problems as BQE:

From QOT to BQE

Assume a database $D$, a domain calculus query $Q = \{ \vec{x} \mid \phi(\vec{x}) \}$, and a value tuple $\vec{c} = (c_1, \ldots, c_n)$. Then $\vec{c} \in Q(D)$ iff $Q'$ evaluates to true in $D$, where

$$Q' = \exists \vec{x}. (\phi(\vec{x}) \land x_1 = c_1 \land \ldots \land x_n = c_n)$$

From QNE to BQE

Assume a database $D$ and a domain calculus query $Q = \{ \vec{x} \mid \phi(\vec{x}) \}$. Then $Q(D) \neq \emptyset$ iff $\exists \vec{x}. \phi(\vec{x})$ evaluates to true in $D$.

From BQE to QNE and QOT $\rightsquigarrow$ trivial.

Relevant Complexity Classes

We recall the inclusions between some fundamental complexity classes:

$$L \subseteq P \subseteq \text{NP} \subseteq \text{PSPACE} \subseteq \text{EXPTIME}$$

- $L$ is the class of all problems solvable in logarithmic space,
- $P$ — in polynomial time,
- $\text{NP}$ — in nondeterministic polynomial time,
- $\text{PSPACE}$ — in polynomial space,
- $\text{EXPTIME}$ — in exponential time.

Complexity Measures for BQE

Combined complexity

The complexity of BQE without any assumptions about the input query $Q$ and database $D$ is called the combined complexity of BQE.

Further measures are obtained by restricting the input:

Data and query complexity

Data complexity of BQE refers to the following decision problem:

INSTANCE: An input database $D$.

QUESTION: Does $Q$ evaluate to true in $D$?

Query complexity of BQE refers to the following decision problem:

INSTANCE: A Boolean query $Q$.

QUESTION: Does $Q$ evaluate to true in $D$?

Complexity of First-order Queries

Theorem (A)

The query complexity and the combined complexity of domain calculus queries is PSPACE-complete (even if we disallow negation and equality atoms). The data complexity is in L (actually, even in a much lower class).

To prove the theorem, we proceed in steps as follows:

1. We provide an algorithm for query evaluation:
   - it shows PSPACE membership for combined complexity (and thus for query complexity as well), and
   - L membership w.r.t. data complexity,

2. We show PSPACE-hardness of query complexity (clearly, the lower bound applies to combined complexity as well).
An algorithm for evaluating FO queries

- We consider an arbitrary FO formula \( \psi \) and a database \( D \).
- W.l.o.g., the formula is of the form
  \[
  \psi = \exists x_1 \forall y_1 \ldots \exists x_n \forall y_n \varphi(x_1, y_1, \ldots, x_n, y_n).
  \]
- Let the active domain \( \text{dom} \) of \( D \) be \( \text{dom} = \{a_1, \ldots, a_m\} \).
- For the evaluation of the formula, we design two procedures \( \text{evaluate}_1 \) and \( \text{evaluate}_2 \), which call each other recursively.
- The algorithm uses global variables \( n \) and \( X = \{x_1, y_1, \ldots, x_n, y_n\} \).

Let us analyse the space usage of our algorithm. We have to store:

1. The input database \( D \) and the formula \( \psi \):
   - do not contribute to the space requirements.
2. The global variables \( X = \{x_1, y_1, \ldots, x_n, y_n\} \):
   - Each variable requires \( O(\log m) \) bits of space. Thus \( X \) needs \( O(n \log m) \) bits. Note that \( X \) requires logarithmic space if \( \psi \) is fixed.
3. A call stack \( S = (S_1, \ldots, S_k) \), where \( k \leq 2n \) and each \( S_j \) stores a state in which a subroutine is called. Clearly, for both subroutines a state \( S_j \) only needs to store the value of \( i \) and the return position in the subroutine:
   - Storing a value \( i \in \{1, \ldots, 2n\} \) requires logarithmic space in the size of \( \psi \) (i.e. \( O(\log n) \)), but only constant space if \( \psi \) is fixed. (The return position requires constant space in both cases.)
   - Hence \( S \) needs \( O(n \log n) \) bits of storage, which is constant if \( \psi \) is fixed.
4. Space for evaluating \( \varphi \) in an assignment:
   - requires a transversal of the parse tree of \( \psi \); space \( O(\log ||\psi||) \) suffices.

Overall we need \( O(n \log m + n \log n + \log ||\psi||) \) bits of storage.

**Proposition**

\( \text{BQE} \in \text{PSPACE} \) w.r.t. combined complexity.
This also implies \( \text{BQE} \in \text{PSPACE} \) w.r.t. query complexity.

For fixed \( \psi \), the space required is \( O(\log m) \), i.e. logarithmic in the data.

**Proposition**

\( \text{BQE} \in \text{L} \) w.r.t. data complexity.

NOTE: Note that \( \text{L} \subseteq \text{P} \). In fact, one can show completeness of \( \text{BQE} \) w.r.t. data complexity for a much lower circuit class \( \text{AC}_0 \subseteq \text{L} \).
The PSPACE lower bound

To prove the PSPACE-hardness result, we first recall quantified Boolean formulae:

**QSAT (QBF)**

**INSTANCE:** An expression $\exists x_1 \forall x_2 \exists x_3 \ldots Q x_n \phi$, where $Q$ is either $\forall$ or $\exists$ and $\phi$ is a Boolean formula in CNF with variables from $\{x_1, x_2, x_3, \ldots, x_n\}$.

**QUESTION:** Is there a truth value for the variable $x_1$ such that for both truth values of $x_2$ there is a truth value for $x_3$ and so on up to $x_n$, such that $\phi$ is satisfied by the overall truth assignment?

**Theorem**

**QSAT** is PSPACE-complete.

**Remark.** A detailed proof is given in the Komplexitätstheorie lecture.

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**Proof of the PSPACE-Hardness of BQE**

The PSPACE-hardness result for Theorem (A) can be shown by a reduction from the QSAT-problem. Let $\psi$ be an arbitrary QBF with

$$\psi = \exists x_1 \forall x_2 \ldots Q x_n \alpha(x_1, \ldots, x_n)$$

where $Q$ is either $\forall$ or $\exists$ and $\alpha$ is a quantifier-free Boolean formula with variables in $\{x_1, \ldots, x_n\}$.

We first define the (fixed) input database $D$ over the predicate symbols $L = \{\text{istrue}, \text{isequal}, \text{not}, \text{or}, \text{and}\}$ with the obvious meaning:

$$D = \{\text{istrue}(1), \text{isequal}(0, 0), \text{isequal}(1, 1), \text{not}(1, 0), \text{not}(0, 1), \text{or}(1, 1), \text{or}(0, 1), \text{or}(0, 1), \text{or}(0, 0), \text{and}(1, 1), \text{and}(1, 0), \text{and}(0, 1), \text{and}(0, 0)\}$$

---

**Proof of the PSPACE-Hardness (continued)**

For each sub-formula $\beta$ of $\alpha$, we define a quantifier-free, first-order formula $T_\beta(z_1, \ldots, z_n, x)$ with the following intended meaning:

if the variables $x_i$ have the truth value $z_i$, then the formula $\beta(x_1, \ldots, x_n)$ evaluates to the truth value $x$.

The formulae $T_\beta(z_1, \ldots, z_n, x)$ can be defined inductively w.r.t. the structure of $\alpha$ as follows:

- $\beta = x_i$ (with $1 \leq i \leq n$): $T_\beta(\bar{z}, x) \equiv \text{isequal}(z_i, x)$
- $\neg \beta'$: $T_\beta(\bar{z}, x) \equiv \exists t_1 T_{\beta'}(\bar{z}, t_1) \land \text{not}(t_1, x)$
- $\beta_1 \land \beta_2$: $T_\beta(\bar{z}, x) \equiv \exists t_1 t_2 T_{\beta_1}(\bar{z}, t_1) \land T_{\beta_2}(\bar{z}, t_2) \land \text{and}(t_1, t_2, x)$
- $\beta_1 \lor \beta_2$: $T_\beta(\bar{z}, x) \equiv \exists t_1 t_2 T_{\beta_1}(\bar{z}, t_1) \land T_{\beta_2}(\bar{z}, t_2) \land \text{or}(t_1, t_2, x)$

**The first-order query $\phi$** is then defined as follows:

$$\phi \equiv \exists x \exists z_1 \forall z_2 \ldots Q z_n T_\alpha(\bar{z}, x) \land \text{istrue}(x)$$

where $Q$ is either $\forall$ or $\exists$ (as in the formula $\psi$).

We claim that this problem reduction is correct, i.e.: The QBF $\psi = \exists x_1 \forall x_2 \ldots Q x_n \alpha(x_1, \ldots, x_n)$ is true $\iff$ the first-order query $\phi \equiv \exists x \exists z_1 \forall z_2 \ldots Q z_n T_\alpha(\bar{z}, x) \land \text{istrue}(x)$ evaluates to true over the database $D$.

The proof is straightforward. It suffices to show by induction on the structure of $\alpha$ that the formulae $T_\beta(z_1, \ldots, z_n, x)$ indeed have the intended meaning.
Complexity of Conjunctive Queries

Recall that conjunctive queries (CQs) are a special case of first-order queries whose only connective is $\land$ and whose only quantifier is $\exists$ (i.e., $\lor$, $\lnot$ and $\forall$ are excluded).

E.g.: $Q = \{ (x) \mid \exists y, z. R(x, y) \land R(y, z) \land P(z, x) \}$

Theorem (B)

The query complexity and the combined complexity of BQE for conjunctive queries is \( \text{NP-complete} \).

Proof

NP-Membership (of the combined complexity). For each variable \( u \) of the query, we guess a domain element to which \( u \) is instantiated. Then we check whether all the resulting ground atoms in the query body exist in \( D \). This check is obviously feasible in polynomial time.

Complexity of Datalog

Theorem (C)

Query evaluation (i.e., the QOT problem) of Datalog has the following complexity:

- \( \text{P-complete w.r.t. data complexity, and} \)
- \( \text{EXPTIME-complete w.r.t. combined and query complexity.} \)

To prove the theorem, we first concentrate on ground Datalog programs:

- A program is ground if it has no variables.
- Such programs are also known as propositional logic programs.
- Note that a ground atom \( R(\text{tim}, \text{bob}) \) can be seen as a propositional variable \( R_{\text{tim,bob}} \).

Proof (continued)

Hardness (of the query complexity). We reduce the NP-complete 3-SAT problem to our problem. For this purpose, we consider the following input database (over a ternary relation symbol \( c \) and a binary relation symbol \( v \)) as fixed:

\[
D = \{ c(1, 1, 1), c(1, 1, 0), c(1, 0, 1), c(1, 0, 0), c(0, 1, 1), c(0, 1, 0), c(0, 0, 1), v(1, 0), v(0, 1) \}
\]

Now let an arbitrary instance of the 3-SAT problem be given through the 3-CNF formula \( \Phi = \bigwedge_{i=1}^{n} I_1 \lor I_2 \lor I_3 \) over the propositional variables \( x_1, \ldots, x_k \). Then we define a conjunctive query \( Q \) as follows:

\[(\exists x_1, \ldots, x_k, \bar{x}_1, \ldots, \bar{x}_k) \quad c(I_1, I_2, I_3) \land \cdots \land c(I_{n_1}, I_{n_2}, I_{n_3}) \land v(x_1, \bar{x}_1) \land \cdots \land v(x_k, \bar{x}_k)\]

where \( I^* = l \) if \( l = x \), and \( I^* = \bar{x} \) if \( l = \neg x \). Moreover, \( \bar{x}_1, \ldots, \bar{x}_k \) are fresh first-order variables. By slight abuse of notation, we thus use \( x_i \) to denote either a propositional atom (in \( \Phi \)) or a first-order variable (in \( Q \)).

It is straightforward to verify that the 3-CNF formula \( \Phi \) is satisfiable \( \iff \) \( Q \) evaluates to true in \( D \).
P-hardness of Ground Datalog

Proof: (Hardness)

- By encoding of a TM. Assume $M = (K \cup \Sigma, \delta, q_{start})$, an input string $I$ and a number of steps $N$, where $N$ is a polynomial of $|I|$.
- We construct in logspace a program $P(M, N)$, a database $DB(I, N)$ and an atom $A$ such that $A \in TP_{P(M, N)}(DB(I, N))$ iff $M$ accepts $I$ in $N$ steps.
- Recall that the transition function $\delta$ of $M$ with a single tape can be represented by a table whose rows are tuples $t = (q_1, \sigma_1, q_2, \sigma_2, d)$. Such a tuple $t$ expresses the following if-then-rule:
  
  - if at some time instant $\tau$ the machine is in state $q_1$, the cursor points to cell number $\pi$, and this cell contains symbol $\sigma_1$,
  - then at instant $\tau + 1$ the machine is in state $q_2$, cell number $\pi$ contains symbol $\sigma_2$, and the cursor points to cell number $\pi + d$.

P-hardness of Ground Datalog: the Database

The construction of the database $DB(I, N)$:

- $\text{symbol}_I[0, 0]$,
- $\text{symbol}_I[0, \pi], \text{ for } 0 < \pi \leq |I|, \text{ where } l_{\pi} = \sigma$
- $\text{symbol}_I[0, \pi], \text{ for } |I| < \pi \leq N$
- $\text{cursor}[0, 0]$,
- $\text{state}_{q_{start}}[0]$.

P-hardness of Ground Datalog: the Rules

- transition rules: for each entry $(q_1, \sigma_1, q_2, \sigma_2, d), 0 \leq \tau < N, 0 \leq \pi < N, 0 \leq \pi + d$.
  
  - $\text{symbol}_I[\tau + 1, \pi] \leftarrow \text{state}_{\tau}[\tau], \text{symbol}_I[\tau, \pi], \text{cursor}[\tau, \pi]$
  - $\text{cursor}[\tau + 1, \pi + d] \leftarrow \text{state}_{\tau}[\tau], \text{symbol}_I[\tau, \pi], \text{cursor}[\tau, \pi]$
  - $\text{state}_{\tau}[\tau + 1] \leftarrow \text{state}_{\tau}[\tau], \text{symbol}_I[\tau, \pi], \text{cursor}[\tau, \pi]$

- inertia rules: where $0 \leq \tau < N, 0 \leq \pi < \pi' \leq N$
  
  - $\text{symbol}_I[\tau + 1, \pi] \leftarrow \text{symbol}_I[\tau, \pi], \text{cursor}[\tau, \pi']$
  - $\text{symbol}_I[\tau + 1, \pi'] \leftarrow \text{symbol}_I[\tau, \pi], \text{cursor}[\tau, \pi]$

- accept rule: for $0 \leq \tau < N$
  
  - $\text{accept} \leftarrow \text{state}_{q_{yes}}[\tau]$
P-hardness of Ground Datalog

The encoding precisely simulates the behaviour of $M$ on input $I$ up to $N$ steps. (This can be formally shown by induction on the time steps.)

- accept $\in T^w_{P(M,N)}(DB(I,N))$ iff $M$ accepts $I$ in $N$ steps.
- The construction is feasible in logarithmic space.
- Note that each rule in $P(M,N)$ has at most 4 atoms. In fact, $P$-hardness applies already for programs with at most 3 atoms in the rules:
  - Simply replace each $A \leftarrow B, C, D$ in $P(M,N)$ by $A \leftarrow B, E$ and $E \leftarrow C, D$, where $E$ is a fresh atom.

Grounding Complexity

Given a program $P$ and a database $DB$, the number of rules in $\text{ground}(P, DB)$ is bounded by

$$|P| \cdot \#\text{consts}(P, DB)^{v_{\text{max}}}$$

- $v_{\text{max}}$ is the maximum number of different variables in any rule $r \in P$
- $\#\text{consts}(P, DB)$ is the number of constants occurring in $P$ and $DB$.
- $\text{ground}(P, DB)$ is polynomial in the size of $DB$.
- Hence, the complexity of propositional logic programming is an upper bound for the data complexity.
- Note that $\text{ground}(P, DB)$ can be exponential in the size of $P$.

Data Complexity of Datalog

### Proposition

*Query evaluation in Datalog is $P$-complete w.r.t. data complexity.*

### Proof: (Membership)

Effective reduction to reasoning over ground Datalog programs is possible. Given a program $P$, a database $DB$, and atom $A$:

- Generate $P' = \text{ground}(P, DB)$, i.e. the set of all ground instances of rules in $P$:
  $$\text{ground}(P, DB) = \bigcup_{r \in P} \text{Ground}(r; P, DB)$$
  
  NB: more details on $\text{Ground}(r; P, DB)$ in Lecture 2.
- Decide whether $A \in T^w_{P'}(DB)$.

### Data Complexity of Datalog: $P$-hardness

#### Proof: Hardness

The $P$-hardness can be shown by writing a simple Datalog *meta-interpreter* for ground programs with at most 3 atoms per rule:

- Represent rules $A_0 \leftarrow A_1, \ldots, A_i$ of such a program $P$, where $0 \leq i \leq 2$, using database facts $(A_0, \ldots, A_i)$ in an $(i+1)$-ary relation $R_i$ on the propositional atoms.
- Then, the program $P$ which is stored this way in a database $DB_{MI}(P)$ can be evaluated by a fixed Datalog program $P_{MI}$ which contains for each relation $R_i$, $0 \leq i \leq 2$, a rule
  $$T(X_0) \leftarrow T(X_1), \ldots, T(X_i), R(X_0, \ldots, X_i).$$
  $$T(x)$$ intuitively means that atom $x$ is true. Then,
  $$A \in T^w_{P'}(DB) \iff T(A) \in T^w_{P_{MI}}(DB_{MI}(P))$$
- $P$-hardness of the data complexity of Datalog is then immediately obtained.
Combined and Query Complexity of Datalog

Proposition

Datalog is EXPTIME-complete w.r.t. query and combined complexity.

Proof

(Membership) Grounding $P$ using $DB$ leads to a propositional program $\text{ground}(P, DB)$ whose size is exponential in the size of $P$ and $DB$. Hence, the query and the combined complexity is in EXPTIME.

(Hardness) We show hardness for query complexity only. Goal: adapt our previous encoding of $TM$ $M$ and input $I$ to obtain a program $P_{\text{dat}}(M, I, N)$ and a fixed database $DB_{\text{dat}}$ to decide acceptance of $M$ on $I$ within $N = 2^m$ steps, where $m = n^k(n = |I|)$ is a polynomial.

Note: We are not allowed to generate an exponentially large program by using exponentially many propositional atoms (the reduction would not be polynomial!)

Query Complexity of Datalog: EXPTIME-hardness

The predicates $\text{Succ}^m$, $\text{First}^m$, and $\text{Last}^m$ are provided.

- The database facts $\text{symbol}_s[0, \pi]$ are readily translated into the Datalog rules

  $$\text{symbol}_s(\mathbf{X}, t) \leftarrow \text{First}^m(\mathbf{X}),$$

  where $t$ represents the position $\pi$.

- Similarly for the facts $\text{cursor}[0, 0]$ and $\text{state}_{\epsilon}[0]$.

- Database facts $\text{symbol}_s[0, \pi]$, where $|I| \leq \pi \leq N$, are translated to the rule

  $$\text{symbol}_s(\mathbf{X}, \mathbf{Y}) \leftarrow \text{First}^m(\mathbf{X}), \leq^m(t, \mathbf{Y})$$

  where $t$ represents the number $|I| + 1$.

Query Complexity of Datalog: EXPTIME-hardness

Ideas for lifting $P(M, N)$ and $DB(I, N)$ to $P_{\text{dat}}(M, I, N)$ and $DB_{\text{dat}}$:

- use the predicates $\text{symbol}_s(X, Y)$, $\text{cursor}(X, Y)$ and $\text{state}_s(X)$ instead of the propositional letters $\text{symbol}_s[X, Y]$, $\text{cursor}[X, Y]$ and $\text{state}_s[X]$ respectively.

- W.l.o.g., let $N$ be of the form $N = 2^m - 1$ for some integer $m \geq 1$.

  The time points $\tau$ and tape positions $\pi$ from 0 to $N$ are encoded in binary, i.e. by $m$-ary tuples $\tau_i = (c_1, \ldots, c_m)$, $c_i \in \{0, 1\}$, $i = 1, \ldots, m$, such that $0 = (0, \ldots, 0)$, $1 = (0, \ldots, 1)$, $N = (1, \ldots, 1)$.

- The functions $\tau + 1$ and $\pi + d$ are realized by means of the successor $\text{Succ}^m$ from a linear order $\leq^m$ on $\{0, 1\}^m$.

- Transition and inertia rules: for realizing $\tau + 1$ and $\pi + d$, use in the body atoms $\text{Succ}^m(X, X')$. For example, the clause

  $$\text{symbol}_s[\tau + 1, \pi] \leftarrow \text{state}_{\epsilon}[:\tau], \text{symbol}_s[\tau, \pi], \text{cursor}[\tau, \pi]$$

  is translated into

  $$\text{symbol}_s(X', Y) \leftarrow \text{state}_{\epsilon}(X), \text{symbol}_s(X, Y), \text{cursor}(X, Y), \text{Succ}^m(X, X').$$

- The translation of the accept rules is straightforward:

  $$\text{accept} \leftarrow \text{state}_{\epsilon}(X).$$
Defining $\text{Succ}^m(X, X')$ and $\leq^m$

- The ground facts $\text{Succ}^1(0, 1)$, $\text{First}^i(0)$, and $\text{Last}^i(1)$ are provided in $DB_{\text{dat}}$.
- For an inductive definition, suppose $\text{Succ}^i(X, Y)$, $\text{First}^i(X)$, and $\text{Last}^i(X)$ tell the successor, the first, and the last element from a linear order $\leq^i$ on $\{0, 1\}^i$, where $X$ and $Y$ have arity $i$.

Then, use rules

$$\text{Succ}^{i+1}(Z, X, Z, Y) \leftarrow \text{Succ}^i(X, Y)$$
$$\text{Succ}^{i+1}(Z, X, Z', Y) \leftarrow \text{Succ}^i(Z, Z'), \text{Last}^i(X), \text{First}^i(Y)$$
$$\text{First}^{i+1}(Z, X) \leftarrow \text{First}^i(Z), \text{First}^i(X)$$
$$\text{Last}^{i+1}(Z, X) \leftarrow \text{Last}^i(Z), \text{Last}^i(X)$$

Defining $\text{Succ}^m(X, X')$ and $\leq^m$

- The ground facts $\text{Succ}^1(0, 1)$, $\text{First}^i(0)$, and $\text{Last}^i(1)$ are provided in $DB_{\text{dat}}$.
- For an inductive definition, suppose $\text{Succ}^i(X, Y)$, $\text{First}^i(X)$, and $\text{Last}^i(X)$ tell the successor, the first, and the last element from a linear order $\leq^i$ on $\{0, 1\}^i$, where $X$ and $Y$ have arity $i$.

Alternatively, use rules

$$\text{Succ}^{i+1}(0, X, 0, Y) \leftarrow \text{Succ}^i(X, Y)$$
$$\text{Succ}^{i+1}(1, X, 1, Y) \leftarrow \text{Succ}^i(X, Y)$$
$$\text{Succ}^{i+1}(0, X, 1, Y) \leftarrow \text{Last}^i(X), \text{First}^i(Y)$$
$$\text{First}^{i+1}(0, X) \leftarrow \text{First}^i(X)$$
$$\text{Last}^{i+1}(1, X) \leftarrow \text{Last}^i(X)$$

The order $\leq^m$ is easily defined from $\text{Succ}^m$ by two clauses

$$\leq^m(X, X) \leftarrow$$
$$\leq^m(X, Y) \leftarrow \text{Succ}^m(X, Z), \leq^m(Z, Y)$$

Combined and Query Complexity of Datalog: Conclusion

- Let $L$ be an arbitrary language in EXPTIME, i.e., there exists a Turing machine $M$ deciding $L$ in exponential time. Then there is a constant $k$ such that $M$ accepts/rejects every input $I$ within $2^{|I|^k}$ steps.
- The program $P_{\text{dat}}(M, I, |I|^k)$ is constructible from $M$ and $I$ in polynomial time (in fact, careful analysis shows feasibility in logarithmic space).
- $\text{accept}$ is in the answer of $P_{\text{dat}}(M, I, |I|^k)$ evaluated over $DB_{\text{dat}} \leftrightarrow M$ accepts input $I$ within $N$ steps.
- Thus the EXPTIME-hardness follows.
Complexity of Datalog with Stratified Negation

**Theorem**

*Reasoning in stratified ground Datalog programs with negation is P-complete. Stratified Datalog with negation is*

- P-complete \(\) w.r.t. data complexity, and
- \(\text{EXPTIME}\)-complete w.r.t. combined and query complexity.

- A ground stratified program \(P\) can be partitioned into disjoint sets \(S_1, \ldots, S_n\) s.t. the semantics of \(P\) is computed by successively computing in polynomial time the fixed-points of the immediate consequence operators \(T_{S_1}, \ldots, T_{S_n}\).

- As with plain Datalog, for programs with variables, the grounding step causes an exponential blow-up.

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**Learning Objectives**

- The BQE, QOT and QNE problems
- The notions of combined, data and query complexity
- The complexity of first-order queries
- The complexity of conjunctive queries
- The complexity of plain and stratified Datalog