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Informatique théorique et applications, tome 21, n° 4 (1987),
p. 419-435

http://www.numdam.org/item?id=ITA_1987__21_4_419_0

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THE DIFFERENCE AND TRUTH-TABLE HIERARCHIES FOR NP^(*)

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Communicated by K. MEHLHORN

Abstract. – Two hierarchies of complexity classes are defined. Both, the difference hierarchy and the truth-table hierarchy for NP are located between NP as bottom class and Δ_2^P . We give examples for complete sets in both hierarchies and investigate their interrelationships. It turns out that the “ ω -jump” of the truth-table hierarchy depends on the encodings of Boolean functions that we use.

Résumé. – Nous définissons deux hiérarchies des classes de complexité. La hiérarchie de différences et la hiérarchie du tableau de vérité pour NP sont localisés entre NP comme la classe de base et Δ_2^P . Nous présentons des exemples d'ensembles complets dans les deux hiérarchies, et nous examinons leurs corrélations. Il apparaît que le « ω -bond » de la hiérarchie du tableau de vérité dépend du chiffrement utilisé des formules de Boole.

1. INTRODUCTION

The class Δ_2^P –NP has received a certain current research interest since sets in this class arise from many NP-complete problems by imposing certain uniqueness or optimality conditions (see [1, 6, 7, 8, 13]). For example, consider the well known NP-complete set CLIQUE, i. e. the set of pairs (G, k) such that G is an undirected graph containing a clique of (at least) k vertices. Then MAXCLIQUE—the problem of deciding whether k is the maximum clique size of G —is in Δ_2^P –NP, provided $\text{NP} \neq \text{coNP}$ (see [9]). In [9] also other NP-complete problems are described where this phenomenon occurs.

(*) Received in June 1986.

This paper is a revised and extended version of the diploma thesis of the first author (see [6]).

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Papadimitriou and Yannakakis [11] introduced a new class \mathbf{D}^P located between \mathbf{NP} and Δ_2^P which seems appropriate to give an exact characterization for some of the problems occurring in $\Delta_2^P - \mathbf{NP}$, since some of those problems, like $\mathbf{MAXCLIQUE}$, turn out to be \mathbf{D}^P -complete. In [10], Papadimitriou also presents some similar new problems that are even complete for Δ_2^P . In [3], [6] and [14] hierarchies have been introduced and investigated independently which are located between \mathbf{NP} as bottom class and Δ_2^P . The union of each of these hierarchies is the Boolean closure of \mathbf{NP} .

In the present paper which should be considered as a revised and extended version of [6] we show that all these hierarchies either coincide or are very closely related to each other. For each level of these hierarchies we give examples of complete sets, and we present conditions causing these hierarchies to “collapse”.

Finally, we study the complexities of the “ ω -jumps” of these hierarchies, and we obtain the somewhat surprising result that the answer depends heavily on the succinctness of the encodings of the Boolean functions that we use. If, for example, the Boolean functions are represented by Boolean circuits then it is very likely that the “ ω -jump” has a higher complexity than in the case when the Boolean functions are represented by Boolean formulas.

2. NOTATION

All our sets are languages over some fixed alphabet Σ , say $\Sigma = \{0, 1\}$. For a string $w \in \Sigma^*$, $|w|$ denotes its length, and for a set $A \subseteq \Sigma^*$, $|A|$ denotes its cardinality. The symmetric difference of two sets A and B is defined by $A \Delta B = (A - B) \cup (B - A)$. For a set $A \subseteq \Sigma^*$, let $\bar{A} = \Sigma^* - A$ be its complement, and for classes \mathbf{A} and \mathbf{B} of sets let $\text{co } \mathbf{A} = \{\bar{A} : A \in \mathbf{A}\}$, $\mathbf{A} \wedge \mathbf{B} = \{A \cap B : A \in \mathbf{A} \text{ and } B \in \mathbf{B}\}$, $\mathbf{A} \vee \mathbf{B} = \{A \cup B : A \in \mathbf{A} \text{ and } B \in \mathbf{B}\}$, and let $\mathbf{BA}(\mathbf{A})$ denote the Boolean algebra generated by \mathbf{A} , i. e. the smallest class that contains \mathbf{A} and is closed under union, intersection and complementation.

Our model of computation is the multi-tape Turing machine. For a Turing machine M , let $L(M)$ denote the set accepted by M , and for an oracle Turing machine and an oracle set A , let $L(M, A)$ be the set accepted by M when using the oracle A . Let $\mathbf{P}(\mathbf{NP})$ be the class of sets acceptable by deterministic (nondeterministic) polynomial-time bounded Turing machines. A Turing machine is said to be polynomial-time bounded if there is a polynomial p such that each computation of M on inputs of size n has length at most $p(n)$. Similarly, let $\mathbf{P}^A(\mathbf{NP}^A)$ be the class of sets acceptable by deterministic (nondeterministic) polynomial-time bounded oracle machines when using the

oracle set A . The class $\cup\{\mathbf{P}^A : A \in \mathbf{NP}\}$ is commonly called $\Delta_2^{\mathbf{P}}$ due to its membership in a more general structure, the polynomial-time hierarchy (see [13]). If $A \in \mathbf{P}^B$ then we also say that A is polynomial-time Turing reducible to B (for short: $A \leq_T^{\mathbf{P}} B$). A more restrictive but more commonly used reducibility is the polynomial-time many-one reducibility: A is polynomial-time many-one reducible to B (for short $A \leq_m^{\mathbf{P}} B$) if there is a function $f: \Sigma^* \rightarrow \Sigma^*$ computable in polynomial time such that $f^{-1}(B) = A$. For a set A and a class \mathbf{C} of sets, A is called \mathbf{C} -hard (w. r. t. $\leq_m^{\mathbf{P}}$) if $B \leq_m^{\mathbf{P}} A$ for each $B \in \mathbf{C}$. Further, A is called \mathbf{C} -complete if A is \mathbf{C} -hard and $A \in \mathbf{C}$. A well known \mathbf{NP} -complete set is SAT, the set of (encodings of) satisfiable Boolean formulas.

Finally, for a set A let $c_A: \Sigma^* \rightarrow \{0, 1\}$ denote its characteristic function.

3. THE DIFFERENCE HIERARCHY

In this section we introduce and investigate a hierarchy of classes located between \mathbf{NP} as the bottom class and $\Delta_2^{\mathbf{P}} = \cup\{\mathbf{P}^A : A \in \mathbf{NP}\}$. The definition is based on the symmetric difference operation and is motivated from the definition of the “*n-r.e.*” sets in recursive function theory (see [12], p. 67). The second level of the difference hierarchy turns out to be Papadimitriou and Yannakakis’ class $\mathbf{D}^{\mathbf{P}}$ [11], and the union of the difference hierarchy is the Boolean algebra generated by the sets in \mathbf{NP} . We show that the difference hierarchy essentially coincides with the Boolean \mathbf{NP} -hierarchy introduced in [14] and with the Boolean hierarchy introduced in [3]. Thus this hierarchy has been introduced and investigated independently in [3], [6] and [14]. Further, we define complete sets for each level of this hierarchy and study conditions under which the hierarchy “collapses”, i.e. it consists only of finitely many levels.

DEFINITION 3.1: The *difference hierarchy* (for \mathbf{NP}) is the sequence of classes $\{\mathbf{DIFF}_k\}_{k \geq 1}$ where $\mathbf{DIFF}_1 = \mathbf{NP}$ and $\mathbf{DIFF}_{k+1} = \{A \Delta B : A \in \mathbf{DIFF}_k \text{ and } B \in \mathbf{NP}\}$ for each $k \geq 1$. Further, let $\mathbf{DH} = \mathbf{DIFF}_1 \cup \mathbf{DIFF}_2 \cup \mathbf{DIFF}_3 \cup \dots$.

Clearly, we have $\mathbf{NP} = \mathbf{DIFF}_1 \subseteq \mathbf{DIFF}_2 \subseteq \dots \subseteq \mathbf{DH} \subseteq \Delta_2^{\mathbf{P}}$. However, it is not known whether any of these inclusions is strict. If any of these inclusions is strict then $\mathbf{NP} \neq \mathbf{coNP}$. Next we summarize some elementary properties of this hierarchy.

PROPOSITION 3.2: For every $k \geq 1$,

- (i) $\text{DIFF}_k \cup \text{co DIFF}_k \subseteq \text{DIFF}_{k+1} \cap \text{co DIFF}_{k+1}$.
- (ii) DIFF_k is closed under polynomial-time many-one reducibility.
- (iii) $\text{DIFF}_k = \text{DIFF}_{k+1}$ implies $\text{DIFF}_k = \text{DH}$.

Proof: (i) Trivial because of $\overline{A \Delta B} = \overline{A} \Delta B$, $A = A \Delta \emptyset$ and $\overline{A} = A \Delta \Sigma^*$.

(ii) Suppose $A \leq_m^p B$ via some polynomial-time computable function f , and suppose that $B \in \text{DIFF}_k$. then there exist sets C_1, \dots, C_k in NP such that $B = C_1 \Delta \dots \Delta C_k$. Now we have

$$A = f^{-1}(B) = f^{-1}(C_1 \Delta \dots \Delta C_k) = f^{-1}(C_1) \Delta \dots \Delta f^{-1}(C_k).$$

Since $f^{-1}(C_i) \in \text{NP}$ for all $i = 1, \dots, k$, we have $A \in \text{DIFF}_k$.

(iii) By induction on n it can easily be seen that $\text{DIFF}_{k+n} = \text{DIFF}_k$ for each $n \geq 1$. ■

It turns out that the difference hierarchy essentially coincides with the Boolean NP -hierarchy introduced in [14] as the sequence of classes $\{\text{C}_k^{\text{NP}}, \text{D}_k^{\text{NP}}\}_{k \geq 1}$ and with the Boolean hierarchy introduced in [3] as the sequence of classes $\{\text{NP}(\mathbf{k})\}_{k \geq 1}$.

DEFINITION 3.3: (i) [14] For every $k \geq 1$,

$$\text{C}_{2k-1}^{\text{NP}} = \text{co NP} \vee \bigvee_{k-1} (\text{NP} \wedge \text{co NP}), \quad \text{D}_{2k-1}^{\text{NP}} = \text{NP} \vee \bigvee_{k-1} (\text{NP} \wedge \text{co NP})$$

$$\text{C}_{2k}^{\text{NP}} = \bigvee_k (\text{NP} \wedge \text{co NP}), \quad \text{D}_{2k}^{\text{NP}} = \text{NP} \vee \text{co NP} \vee \bigvee_{k-1} (\text{NP} \wedge \text{co NP})$$

(ii) [3] $\text{NP}(\mathbf{1}) = \text{NP}$ and, for every $k \geq 1$,

$$\text{NP}(\mathbf{2k}) = \text{NP}(\mathbf{2k-1}) \wedge \text{co NP}, \quad \text{NP}(\mathbf{2k+1}) = \text{NP}(\mathbf{2k}) \vee \text{NP}.$$

To prove the relationships between the Boolean NP -hierarchy, the difference hierarchy and the Boolean hierarchy we need the following definition.

DEFINITION 3.4: Let f be a k -ary Boolean function. We define: $A \in f(\text{NP})$ iff there exist $B_1, \dots, B_k \in \text{NP}$ such that $c_A(x) = f(c_{B_1}(x), \dots, c_{B_k}(x))$ for all $x \in \Sigma^*$.

It has been shown that for every Boolean function f the class $f(\text{NP})$ coincides with one of the classes C_k^{NP} or D_k^{NP} . To answer the question of which C_k^{NP} or D_k^{NP} is equal to $f(\text{NP})$ for a given Boolean function f the non-numerical measure c for Boolean functions has been introduced. Let \sqsubseteq be the initial word relation, i. e. $u \sqsubseteq w$ iff there exists a v such that $uv = w$, for

$u, w \in \Sigma^*$. Further, for $a_1, \dots, a_k, b_1, \dots, b_k \in \{0, 1\}$ we define $(a_1, \dots, a_k) \leq (b_1, \dots, b_k)$ iff $a_1 \leq b_1, \dots, a_k \leq b_k$. Finally, set $\underline{a}_i = (a_{i1}, \dots, a_{ik})$, and define $c(f) = \max \{ \{ f(\underline{a}_1) f(\underline{a}_2) \dots f(\underline{a}_r) : r \geq 1, f(\underline{a}_i) \neq f(\underline{a}_{i+1}) \text{ and } \underline{a}_1 \leq \underline{a}_2 \leq \dots \leq \underline{a}_r \} \}$ (it is obvious that this maximum exists) and $m(f) = |c(f)| - 1$.

THEOREM 3.5 [14]: *For every Boolean function f ,*

$$f(\mathbf{NP}) = \begin{cases} \mathbf{C}_{m(f)}^{\mathbf{NP}} & \text{if } c(f) \text{ ends with } 0 \\ \mathbf{D}_{m(f)}^{\mathbf{NP}} & \text{if } c(f) \text{ ends with } 1. \blacksquare \end{cases}$$

Using the above theorem it is not hard to prove.

THEOREM 3.6 [14]: *For every $k \geq 1$,*

- (i) $\mathbf{C}_{k+1}^{\mathbf{NP}} = \mathbf{D}_k^{\mathbf{NP}} \wedge \mathbf{co NP}$,
- (ii) $\mathbf{D}_{k+1}^{\mathbf{NP}} = \mathbf{C}_k^{\mathbf{NP}} \vee \mathbf{NP}$,
- (iii) $\mathbf{C}_k^{\mathbf{NP}} = \mathbf{co D}_k^{\mathbf{NP}}$,
- (iv) $\mathbf{C}_k^{\mathbf{NP}} \cup \mathbf{D}_k^{\mathbf{NP}} \subseteq \mathbf{C}_{k+1}^{\mathbf{NP}} \cap \mathbf{D}_{k+1}^{\mathbf{NP}}$,
- (v) $\bigcup_{k \geq 1} \mathbf{C}_k^{\mathbf{NP}} = \bigcup_{k \geq 1} \mathbf{D}_k^{\mathbf{NP}} = \mathbf{BA}(\mathbf{NP})$. \blacksquare

Note that the above results are proved in [14] not only for **NP** but also for all classes which are closed under union and intersection.

From Theorem 3.6 (i) and (ii) we obtain immediately the relationships between the Boolean **NP**-hierarchy from [14] and the Boolean hierarchy from [3].

COROLLARY 3.7: *For every $k \geq 1$,*

- (i) $\mathbf{NP}(2k-1) = \mathbf{D}_{2k-1}^{\mathbf{NP}}$,
- (ii) $\mathbf{NP}(2k) = \mathbf{C}_{2k}^{\mathbf{NP}}$. \blacksquare

Theorem 3.5 enables us to establish the result on the relationship between the difference hierarchy and the Boolean **NP**-hierarchy.

THEOREM 3.8: *For every $k \geq 1$,*

- (i) $\mathbf{DIFF}_{2k-1} = \mathbf{D}_{2k-1}^{\mathbf{NP}}$,
- (ii) $\mathbf{DIFF}_{2k} = \mathbf{C}_{2k}^{\mathbf{NP}}$.

Proof: Obviously, $\mathbf{DIFF}_k = \mathbf{d}_k(\mathbf{NP})$ where

$$d_k(x_1, x_2, \dots, x_k) = x_1 \oplus x_2 \oplus \dots \oplus x_k,$$

where \oplus stands for addition modulo 2. Furthermore, because of

$$d_k(0, 0, 0, \dots, 0) d_k(1, 0, 0, \dots, 0) d_k(1, 1, 0, \dots, 0) \dots d_k(1, 1, 1, \dots, 1) = 0101 \dots$$

and since

$$(0, 0, 0, \dots, 0) \leq (1, 0, 0, \dots, 0) \leq (1, 1, 0, \dots, 0) \leq \dots \leq (1, 1, 1, \dots, 1)$$

is a chain of maximum length we have $c(d_k) = 0101\dots$. By Theorem 3.5 we have, for $k \geq 1$,

$$\mathbf{d}_{2^k-1}(\mathbf{NP}) = \mathbf{D}_{2^k-1}^{\mathbf{NP}} \quad \text{and} \quad \mathbf{d}_{2^k}(\mathbf{NP}) = \mathbf{C}_{2^k}^{\mathbf{NP}}. \blacksquare$$

COROLLARY 3.9: (i) For every $k \geq 1$, $\mathbf{DIFF}_k = \mathbf{NP}(k)$.

(ii) The sequence $\{\mathbf{DIFF}_k, \mathbf{coDIFF}_k\}_{k \geq 1}$ is the Boolean NP-hierarchy.

(iii) $\mathbf{DH} = \mathbf{BA}(\mathbf{NP})$.

(iv) For every $k \geq 1$, $\mathbf{DIFF}_k = \mathbf{coDIFF}_k$ implies $\mathbf{DIFF}_k = \mathbf{DH}$.

Proof: Only (iv) deserves a proof. Assume $\mathbf{DIFF}_k = \mathbf{coDIFF}_k$. Let $k = 2m$. We conclude

$$\begin{aligned} \mathbf{coDIFF}_{k+1} &= \mathbf{coDIFF}_{2m+1} = \mathbf{C}_{2m+1}^{\mathbf{NP}} = \bigvee_m (\mathbf{NP} \wedge \mathbf{coNP}) \vee \mathbf{coNP} = \mathbf{C}_{2m}^{\mathbf{NP}} \vee \mathbf{coNP} \\ &= \mathbf{DIFF}_k \vee \mathbf{coNP} = \mathbf{coDIFF}_k \vee \mathbf{coNP} = \mathbf{D}_{2m}^{\mathbf{NP}} \vee \mathbf{coNP} \\ &= \bigvee_{m-1} (\mathbf{NP} \wedge \mathbf{coNP}) \vee \mathbf{NP} \vee \mathbf{coNP} \vee \mathbf{coNP} = \mathbf{D}_{2m}^{\mathbf{NP}} = \mathbf{coDIFF}_k. \end{aligned}$$

Consequently, $\mathbf{DIFF}_k = \mathbf{DIFF}_{k+1}$, and by Proposition 3.2 (iii) we have $\mathbf{DIFF}_k = \mathbf{DH}$. The case $k = 2m + 1$ is treated analogously. \blacksquare

Now we consider sets which are complete for every level of the difference hierarchy.

DEFINITION 3.10: Let L be an arbitrary language and let $k \geq 1$. Define

$$L^{k-\Delta} = \{(x_1, \dots, x_k) : |\{i : x_i \in L\}| \text{ is odd}\} \quad \text{and} \quad L^{\omega-\Delta} = \bigcup_{k \geq 1} L^{k-\Delta}.$$

THEOREM 3.11: (i) For each $k \geq 1$, $\mathbf{SAT}^{k-\Delta}$ is \mathbf{DIFF}_k -complete.

(ii) $\mathbf{SAT}^{\omega-\Delta}$ is \mathbf{DH} -hard.

Proof: (i) First note that

$$\mathbf{SAT}^{k-\Delta} = \bigwedge_{i=1}^k (\Sigma^* \times \Sigma^* \times \dots \times \Sigma^* \times \mathbf{SAT} \times \Sigma^* \times \dots \times \Sigma^*)$$

where \mathbf{SAT} is occurring at position i .

Since all the sets in parenthesis are in \mathbf{NP} it follows that $\mathbf{SAT}^{k-\Delta}$ is in \mathbf{DIFF}_k . Now let A be in \mathbf{DIFF}_k . Then there exist sets A_1, \dots, A_k in \mathbf{NP} and

polynomial-time computable functions f_1, \dots, f_k such that $A = A_1 \Delta \dots \Delta A_k$ and $f_i^{-1}(\text{SAT}) = A_i$ for $i = 1, \dots, k$. Consequently,

$$\begin{aligned} x \in A & \text{ iff } x \in A_1 \Delta \dots \Delta A_k \\ & \text{ iff } |\{i: x \in A_i\}| \text{ is odd} \\ & \text{ iff } |\{i: f_i(x) \in \text{SAT}\}| \text{ is odd} \\ & \text{ iff } (f_1(x), \dots, f_k(x)) \in \text{SAT}^{k-\Delta}. \end{aligned}$$

(ii) follows from (i). ■

Similar DIFF_k -complete sets have been considered independently in [3]. Note that $\text{SAT}^{2-\Delta}$ is essentially the set SAT-UNSAT which has been defined and shown to be D^{P} -complete in [11]. In that paper also some other D^{P} -complete sets have been exhibited. In [16] several natural problems are shown to be complete for every level of the Boolean NP-hierarchy (and thus also for every level of the difference hierarchy). For example, the set $\{(G, n_1, \dots, n_k): \text{the chromatic number of the graph } G \text{ is in } \{n_1, \dots, n_k\}\}$ is complete in $\text{C}_{2k}^{\text{NP}} = \text{co DIFF}_{2k}$.

It is not likely that $\text{SAT}^{\omega-\Delta}$ is in DH, and therefore DH-complete. Moreover, it is even not likely that DH has any complete set. We consider $\text{SAT}^{\omega-\Delta}$ again in Section 5.

COROLLARY 3.12: *The difference hierarchy is finite if and only if DH has a complete set.*

Proof: The forward direction follows from Theorem 3.11 (i). For the backward direction let A be a DH-complete set. Then $A \in \text{DIFF}_k$ for some $k \geq 1$, and by Proposition 3.2 (ii) we have $\text{DH} = \text{DIFF}_k$, hence the difference hierarchy is finite. ■

4. THE TRUTH-TABLE HIERARCHY

In what follows we use a fixed natural, uncontrived encoding of all Boolean circuits with \wedge, \vee and \neg gates. If z is such a encoding of a Boolean circuit then h_z denotes the Boolean function realized by this circuit.

From several equivalent definitions for the polynomial-time truth-table reducibility given in [8] and [2] we choose the following.

DEFINITION 4.1: A set A is polynomial-time truth-table reducible to a set B ($A \leq_{\text{tt}}^{\text{P}} B$) iff there exists a polynomial-time computable function f such that $c_A(x) = h_z(c_B(x_1), \dots, c_B(x_m))$ for all $x \in \Sigma^*$ where $f(x) = (z, x_1, \dots, x_m)$ and

z is the encoding of a Boolean circuit with \wedge , \vee and \neg gates. Note that m depends on x .

(ii) A set A is polynomial-time k -bounded truth-table reducible to B ($A \leq_{k-ut}^P B$) iff $A \leq_{tt}^P B$ via a function f having at most $(k+1)$ -tuples as values. Equivalently, $A \leq_{k-ut}^P B$ iff there exist polynomial-time computable functions g, f_1, \dots, f_k such that $c_A(x) = h_g(x)(c_B(f_1(x)), \dots, c_B(f_k(x)))$ for all $x \in \Sigma^*$ where $g(x)$ is the encoding of a Boolean circuit with \wedge , \vee and \neg gates.

(iii) A set A is polynomial-time bounded truth-table reducible to B ($A \leq_{bt}^P B$) iff $A \leq_{k-ut}^P B$ for some $k \geq 1$.

(iv) For a class of sets C and a reducibility α let

$$\alpha(C) = \{A: \text{there exists a set } B \in C \text{ such that } A \alpha B\}.$$

Note that in [2] a different terminology is used and \leq_{tt}^P is denoted by $\leq_{P.UNIV.ALL}$.

In the preceding definition of \leq_{tt}^P , \leq_{k-ut}^P ($k \geq 1$) and \leq_{bt}^P Boolean functions are represented by Boolean circuits. We also consider the reducibilities \leq_{bf}^P , \leq_{k-bf}^P ($k \geq 1$) and \leq_{tt}^P ($\leq_{f_{tt}}^P$, $\leq_{k-f_{tt}}^P$ ($k \geq 1$) and $\leq_{bf_{tt}}^P$) which are defined in the same way as \leq_{tt}^P , \leq_{k-ut}^P ($k \geq 1$) and \leq_{bt}^P but using Boolean formulas with operations \wedge , \vee and \neg (full truth-tables) instead of Boolean circuits with \wedge , \vee and \neg gates. The following relationships between these reducibilities can easily be proved.

PROPOSITION 4.2: *Let A, B be arbitrary sets, and let $k \geq 1$.*

(i) $A \leq_{f_{tt}}^P B$ implies $A \leq_{bf}^P B$, and $A \leq_{bf}^P B$ implies $A \leq_{tt}^P B$. Consequently,

$$\leq_{f_{tt}}^P(\text{NP}) \subseteq \leq_{bf}^P(\text{NP}) \subseteq \leq_{tt}^P(\text{NP})^{(1)}.$$

(ii) $A \leq_{k-f_{tt}}^P B$ iff $A \leq_{k-bf}^P B$ iff $A \leq_{k-ut}^P B$. Consequently,

$$\leq_{k-f_{tt}}^P(\text{NP}) = \leq_{k-bf}^P(\text{NP}) = \leq_{k-ut}^P(\text{NP}).$$

⁽¹⁾ Note added in proof: Though we had conjectured that $\leq_{bf}^P(\text{NP}) \neq \leq_{tt}^P(\text{NP})$ it turned out recently that these classes coincide. In L. A. HEMACHANDRA, *The Strong Exponential Hierarchy Collapses*, Proc. 19th A.C.M. S.T.O.C., 1987, it is proved that $\leq_{tt}^P(\text{NP})$ is included in the class $\text{P}^{\text{NP}} [O(\log n)]$ of all sets which are polynomial-time computable with $O(\log n)$ queries to an NP-oracle. But the classes $\text{P}^{\text{NP}} [O(\log n)]$ and $\leq_{bf}^P(\text{NP})$ coincide because they have the common complete set ODD CLIQUE. The latter is proved in [16] and M. W. KRENTTEL, *The Complexity of Optimization Problems*, Proc. 18th A.C.M. S.T.O.C., 1986, pp. 69-76. For a further proof of the coincidence of $\leq_{bf}^P(\text{NP})$ and $\leq_{tt}^P(\text{NP})$ see the very recent paper R. J. BEIGEL, *Bounded Queries to SAT and the Boolean Hierarchy*, manuscript, 1987.

(iii) $A \leq_{bftt}^P B$ iff $A \leq_{bbf}^P B$ iff $A \leq_{btt}^P B$. Consequently,

$$\leq_{bftt}^P(\mathbf{NP}) = \leq_{bbf}^P(\mathbf{NP}) = \leq_{btt}^P(\mathbf{NP}).$$

Proof⁽¹⁾ : Every full truth-table can be converted in polynomial time into an equivalent Boolean formula, and every Boolean formula can be converted in polynomial time into an equivalent Boolean circuit. The converse is also true if the number of Boolean variables or the number of input nodes of the Boolean circuits, resp., is bounded to a fixed number $k \geq 1$. Of course, the polynomial time bound depends on k . ■

Clearly, for every class C of languages we have

$$\begin{aligned} C \subseteq \leq_{1-tt}^P(C) \subseteq \leq_{2-tt}^P(C) \subseteq \dots \\ \subseteq \leq_{btt}^P(C) \subseteq \leq_{ftt}^P(C) \subseteq \leq_{bf}^P(C) \subseteq \leq_{tt}^P(C) \subseteq \leq_T^P(C). \end{aligned}$$

We call the sequence \mathbf{NP} , $\leq_{1-tt}^P(\mathbf{NP})$, $\leq_{2-tt}^P(\mathbf{NP})$, ... the (bounded) *truth-table hierarchy* for \mathbf{NP} .

PROPOSITION 4.3: For every $k \geq 1$,

- (i) $\leq_{k-tt}^P(\mathbf{NP}) = \mathbf{co} \leq_{k-tt}^P(\mathbf{NP})$.
- (ii) $\leq_{k-tt}^P(\mathbf{NP})$ is closed under polynomial-time many-one reducibility.

Proof: (i) is obvious; (ii) is proved as Proposition 3.2 (ii). ■

Like the difference hierarchy for \mathbf{NP} , this hierarchy is located between \mathbf{NP} as bottom class and Δ_2^P . Moreover, there exist strong relationships between these hierarchies. To establish our main result in this direction (Theorem 4.5) we need a characterization of polynomial-time k -bounded truth-table reducibility in terms of the Boolean algebra generated by certain sets. The following theorem is a first step in this direction.

THEOREM 4.4: Let $k \geq 1$, and let A and B be arbitrary sets. The following statements are equivalent:

- (i) $A \leq_{k-tt}^P B$
- (ii) there exist sets Q_1, \dots, Q_k which are \leq_m^P -reducible to B such that

$$A \in \mathbf{BA}(P \cup \{Q_1, \dots, Q_k\}),$$

(iii) there exist sets Q_1, \dots, Q_k which are \leq_m^P -reducible to B and sets $P_I \in P$ for every $I \subseteq \{1, \dots, k\}$ such that

$$A = \bigcup_{I \subseteq \{1, \dots, k\}} (P_I \cap \bigcap_{i \in I} Q_i \cap \bigcap_{i \notin I} \bar{Q}_i)$$

Proof: Suppose $A \leq_{k-IT}^P B$. By definition there exist polynomial-time computable functions g, f_1, \dots, f_k such that

$$c_A(x) = h_g(x)(c_B(f_1(x)), \dots, c_B(f_k(x)))$$

for all $x \in \Sigma^*$ where $g(x)$ is the encoding of a Boolean circuit with \wedge, \vee and \neg gates. It is an obvious fact from Boolean function theory that for every k -ary Boolean function h with the variables x_1, \dots, x_k ,

$$\begin{aligned} h(x_1, \dots, x_k) &= \bigvee_{a_1, \dots, a_k \in \{0, 1\}} (h(a_1, \dots, a_k) \wedge \bigvee_{a_i=1} x_i \wedge \bigvee_{a_i=0} \neg x_i) \\ &= \bigvee_{I \subseteq \{1, \dots, k\}} (h(c_I(1), \dots, c_I(k)) \wedge \bigvee_{i \in I} x_i \wedge \bigvee_{i \notin I} \neg x_i). \end{aligned}$$

Consequently,

$$c_A(x) = \bigvee_{I \subseteq \{1, \dots, k\}} (h_g(x)(c_I(1), \dots, c_I(k)) \wedge \bigvee_{i \in I} c_B(f_i(x)) \wedge \bigvee_{i \notin I} \neg c_B(f_i(x))).$$

Defining $Q_i = f_i^{-1}(B)$ for $i = 1, \dots, k$ and

$$P_I = \{x: h_g(x)(c_I(1), \dots, c_I(k)) = 1\} \text{ for } I \subseteq \{1, \dots, k\}$$

we obtain $Q_i \leq_m^P B$, $P_I \in P$ and

$$c_A(x) = \bigvee_{I \subseteq \{1, \dots, k\}} (c_{P_I}(x) \wedge \bigvee_{i \in I} c_{Q_i}(x) \wedge \bigvee_{i \notin I} \neg c_{Q_i}(x))$$

i. e.

$$A = \bigcup_{I \subseteq \{1, \dots, k\}} (P_I \cap \bigcap_{i \in I} Q_i \cap \bigcap_{i \notin I} \bar{Q}_i).$$

The other two directions of the proof are obvious. ■

The preceding theorem gives an equivalent condition for the \leq_{k-IT}^P -reduction from a set A to a specific set B . We obtain a stronger result if we consider the \leq_{k-IT}^P -reduction from A to an arbitrary NP-set.

THEOREM 4.5: *For every $k \geq 1$ there exists a $(k+1)$ -ary Boolean function h_k such that the following statements are equivalent:*

- (i) $A \in \leq_{k-IT}^P(\mathbf{NP})$,
- (ii) *there exist an $m \geq 0$, an $(m+k)$ -ary Boolean function f , sets $A_1, \dots, A_m \in P$ and sets $B_1, \dots, B_k \in \mathbf{NP}$ such that*

$$c_A(x) = f(c_{A_1}(x), \dots, c_{A_m}(x), c_{B_1}(x), \dots, c_{B_k}(x)),$$

(iii) There exist a set $A_1 \in \mathbf{P}$ and sets $B_1, \dots, B_k \in \mathbf{NP}$ such that

$$c_A(x) = h_k(c_{A_1}(x), c_{B_1}(x), \dots, c_{B_k}(x)).$$

Proof: The proof of Theorem 2 in [14] shows that, for fixed $k \geq 1$, there exists polynomial-time computable functions a_1, \dots, a_k such that for every k -ary Boolean function h_z (that is the Boolean function described by the Boolean formula z) and every D_1, \dots, D_k :

- $h_{a_i(z)}$ is a k -ary monotonic Boolean function,
- $h_z(c_{D_1}(x), \dots, c_{D_k}(x)) = \begin{cases} f_k(c_{B_1}(x), \dots, c_{B_k}(x)) & \text{if } c(h_z) \text{ ends with } 0 \\ g_k(c_{B_1}(x), \dots, c_{B_k}(x)) & \text{if } c(h_z) \text{ ends with } 1, \end{cases}$

where $c_{B_i}(x) = h_{a_i(z)}(c_{D_1}(x), \dots, c_{D_k}(x))$, for $i = 1, \dots, k$,

$$f_k(\alpha_1, \dots, \alpha_k) = \begin{cases} \bigvee_{i=1}^{(k-1)/2} (\alpha_{2i-1} \wedge \neg \alpha_{2i}) \vee \neg \alpha_k & \text{if } k \text{ is odd} \\ \bigvee_{i=1}^{k/2} (\alpha_{2i-1} \wedge \neg \alpha_{2i}) & \text{if } k \text{ is even,} \end{cases}$$

$$g_k(\alpha_1, \dots, \alpha_k) = \begin{cases} \bigvee_{i=1}^{(k-1)/2} (\alpha_{2i-1} \wedge \neg \alpha_{2i}) \vee \alpha_k & \text{if } k \text{ is odd} \\ \bigvee_{i=1}^{(k-2)/2} (\alpha_{2i-1} \wedge \neg \alpha_{2i}) \vee \alpha_{k-1} \vee \neg \alpha_k & \text{if } k \text{ is even.} \end{cases}$$

Now let $A \in \leq_{k-\text{tt}}^{\mathbf{P}}(\mathbf{NP})$. By definition there exist polynomial-time computable functions r, s_1, \dots, s_k and a set $B \in \mathbf{NP}$ such that

$$c_A(x) = h_{r(x)}(c_B(s_1(x)), \dots, c_B(s_k(x))).$$

Obviously, the sets $D_i = s_i^{-1}(B)$ are also in \mathbf{NP} ($i = 1, \dots, k$) and we obtain

$$c_A(x) = h_{r(x)}(c_{D_1}(x), \dots, c_{D_k}(x)) = \begin{cases} f_k(c_{B_1}(x), \dots, c_{B_k}(x)) & \text{if } c(h_{r(x)}) \text{ ends with } 0 \\ g_k(c_{B_1}(x), \dots, c_{B_k}(x)) & \text{if } c(h_{r(x)}) \text{ ends with } 1, \end{cases}$$

where $c_{B_i}(x) = h_{a_i(r(x))}(c_{D_1}(x), \dots, c_{D_k}(x))$, for $i = 1, \dots, k$.

Since the Boolean functions $h_{a_i(r(x))}$ are monotonic and D_1, \dots, D_k are in \mathbf{NP} , the sets B_1, \dots, B_k are also in \mathbf{NP} .

Furthermore, the set $A_1 = \{x : c(h_{r(x)}) \text{ ends with } 0\}$ is in \mathbf{P} , because all $h_{r(x)}$ are k -ary Boolean functions.

Defining $h_k(\alpha_0, \alpha_1, \dots, \alpha_k) = (\alpha_0 \wedge f_k(\alpha_1, \dots, \alpha_k)) \vee (\neg \alpha_0 \wedge g_k(\alpha_1, \dots, \alpha_k))$ we obtain

$$\begin{aligned} c_A(x) &= (c_{A_1}(x) \wedge f_k(c_{B_1}(x), \dots, c_{B_k}(x))) \vee (\neg c_{A_1}(x) \wedge g_k(c_{B_1}(x), \dots, c_{B_k}(x))) \\ &= h_k(c_{A_1}(x), c_{B_1}(x), \dots, c_{B_k}(x)). \blacksquare \end{aligned}$$

Note that in the above theorem we can replace **NP** by any class of sets which is closed under \leq_{bf}^P -reducibility via Boolean formulas with operations \vee and \wedge .

COROLLARY 4.6: (i) For every $k \geq 1$,

$$\mathbf{DIFF}_k \cup \mathbf{co\,DIFF}_k \subseteq \leq_{k-itt}^P(\mathbf{NP}) \subseteq \mathbf{DIFF}_{k+1} \cap \mathbf{co\,DIFF}_{k+1}.$$

$$(ii) \leq_{btt}^P(\mathbf{NP}) = \mathbf{BA}(\mathbf{NP}).$$

(iii) For every $k \geq 1$,

$$\leq_{k-itt}^P(\mathbf{NP}) = \leq_{(k+1)-itt}^P(\mathbf{NP}) \text{ implies } \leq_{k-itt}^P(\mathbf{NP}) = \leq_{btt}^P(\mathbf{NP}).$$

(iv) If the truth-table hierarchy is infinite then $\leq_{btt}^P(\mathbf{NP})$ is strictly included in $\leq_{itt}^P(\mathbf{NP})$.

Proof: (i) It is obvious that $\mathbf{DIFF}_k \cup \mathbf{co\,DIFF}_k \subseteq \leq_{k-itt}^P(\mathbf{NP})$.

Let $A \in \leq_{k-itt}^P(\mathbf{NP})$. By Proposition 4.3 (i) we have also $\bar{A} \in \leq_{k-itt}^P(\mathbf{NP})$, and by Theorem 4.5 we have $A, \bar{A} \in h_k(\mathbf{NP})$. Since h_k is $(k+1)$ -ary we can conclude $m(h_k) \leq k+1$. By Theorem 3.5 there holds $h_k(\mathbf{NP}) \subseteq \mathbf{C}_{k+1}^{\mathbf{NP}}$ or $h_k(\mathbf{NP}) \subseteq \mathbf{D}_{k+1}^{\mathbf{NP}}$. Consequently, $A, \bar{A} \in \mathbf{C}_{k+1}^{\mathbf{NP}}$ or $A, \bar{A} \in \mathbf{D}_{k+1}^{\mathbf{NP}}$. From $\mathbf{D}_{k+1}^{\mathbf{NP}} = \mathbf{co\,C}_{k+1}^{\mathbf{NP}}$ [Theorem 3.6 (iii)] we conclude $A \in \mathbf{C}_{k+1}^{\mathbf{NP}} \cap \mathbf{D}_{k+1}^{\mathbf{NP}}$.

(ii) is an immediate consequence of (i) and Corollary 3.9 (ii).

(iii) By (i) the equality $\leq_{k-itt}^P(\mathbf{NP}) = \leq_{(k+1)-itt}^P(\mathbf{NP})$ implies $\leq_{k-itt}^P(\mathbf{NP}) = \mathbf{DIFF}_{k+1} = \mathbf{co\,DIFF}_{k+1}$. From Corollary 3.9 (iv) we conclude $\mathbf{DIFF}_{k+1} = \mathbf{DH}$, and Corollary 3.9 (iii) and Corollary 4.6 (ii) yield $\leq_{k-itt}^P(\mathbf{NP}) = \mathbf{DIFF}_{k+1} = \mathbf{DH} = \mathbf{BA}(\mathbf{NP}) = \leq_{btt}^P(\mathbf{NP})$.

(IV) Obviously, $\mathbf{SAT}^{\omega-\Delta} \subseteq \leq_{itt}^P \mathbf{SAT}$. Therefore, if $\leq_{btt}^P(\mathbf{NP}) = \leq_{itt}^P(\mathbf{NP})$ then $\mathbf{SAT}^{\omega-\Delta} \in \leq_{btt}^P(\mathbf{NP}) = \mathbf{BA}(\mathbf{NP}) = \mathbf{DH}$. By Theorem 3.11 (ii) $\mathbf{SAT}^{\omega-\Delta}$ is **DH**-complete. Using Corollary 3.12 it follows that the difference hierarchy, and consequently also the truth-table hierarchy, is finite. \blacksquare

Thus the union of the difference hierarchy and the union of the truth-table hierarchy are the same. The analogous fact in recursive function theory has already been observed (see [12]).

Next we define a sequence of sets which will subsequently be shown to be complete for the several levels of the truth-table hierarchy.

DEFINITION 4.7: For each set L and $k \geq 1$ define

- (i) $L^{k-tt} = \{ (z, x_1, \dots, x_k) : z \text{ encodes a Boolean circuit with } k \text{ input nodes and } h_z(c_L(x_1), \dots, c_L(x_k)) = 1 \}$,
- (ii) $L^{\omega-tt} = \bigcup_{k \geq 1} L^{k-tt}$,
- (iii) $L^{k-bf} = \{ (z, x_1, \dots, x_k) : z \text{ encodes a Boolean formula with } k \text{ variables and } h_z(c_L(x_1), \dots, c_L(x_k)) = 1 \}$,
- (iv) $L^{\omega-bf} = \bigcup_{k \geq 1} L^{k-bf}$,
- (v) $L^{k-ftt} = \{ (z, x_1, \dots, x_k) : z \text{ encodes a full truth-table with } k \text{ variables and } h_z(c_L(x_1), \dots, c_L(x_k)) = 1 \}$,
- (vi) $L^{\omega-ftt} = \bigcup_{k \geq 1} L^{k-ftt}$.

PROPOSITION 4.8: For each set L and $k \geq 1$,

- (i) $L^{k-ftt} \equiv_m^P L^{k-bf} \equiv_m^P L^{k-tt}$.
- (ii) $L^{\omega-ftt} \leq_m^P L^{\omega-bf} \leq_m^P L^{\omega-tt}$.

Proof: Use the same arguments as for Proposition 4.2 ■

We do not know whether $L^{\omega-ftt} \equiv_m^P L^{\omega-bf}$ or $L^{\omega-bf} \equiv_m^P L^{\omega-tt}$.

THEOREM 4.9: (i) For each $k \geq 1$,

SAT^{k-ftt} , SAT^{k-bf} and SAT^{k-tt} are $\leq_{k-tt}^P(\mathbf{NP})$ -complete.

(ii) $SAT^{\omega-ftt}$, $SAT^{\omega-bf}$ and $SAT^{\omega-tt}$ are $\mathbf{BA}(\mathbf{NP})$ -hard.

Proof: (i) Because of Proposition 4.8 we restrict ourselves to SAT^{k-tt} . Clearly, $SAT^{k-tt} \leq_{k-tt}^P \mathbf{SAT}$. Thus, $SAT^{k-tt} \in \leq_{k-tt}^P(\mathbf{NP})$.

Suppose $A \in \leq_{k-tt}^P(\mathbf{NP})$. Then there exist a $B \in \mathbf{NP}$ and polynomial-time computable functions g, f_1, \dots, f_k such that

$$c_A(x) = h_{g(x)}(c_B(f_1(x)), \dots, c_B(f_k(x))).$$

Because of $B \in \mathbf{NP}$ there exists a polynomial-time computable function g' such that $c_B(x) = c_{\mathbf{SAT}}(g'(x))$ for all $x \in \Sigma^*$. Consequently,

$$\begin{aligned} x \in A \text{ iff } & h_{g(x)}(c_B(f_1(x)), \dots, c_B(f_k(x))) = 1 \\ & \text{iff } h_{g(x)}(c_{\mathbf{SAT}}(g'(f_1(x))), \dots, c_{\mathbf{SAT}}(g'(f_k(x)))) = 1 \end{aligned}$$

iff $(g(x), g'(f_1(x)), \dots, g'(f_k(x))) \in \text{SAT}^{k-t}$.

(ii) follows from (i). ■

5. BEYOND THE HIERARCHIES

In this section we consider the exact complexity of the sets $\text{SAT}^{\omega-\Delta}$, $\text{SAT}^{\omega-ft}$, $\text{SAT}^{\omega-bf}$ and $\text{SAT}^{\omega-t}$ defined in sections 3 and 4. That is, we are interested in the complexity of the different natural “ ω -jumps” of our hierarchies. Note that one answer for the analogous question concerning the polynomial-time hierarchy is that the ω -jump B_ω defined in [13] is **PSPACE**-complete. However, in [5] another ω -jump K^ω of the polynomial-time hierarchy has been considered, and it is an open question whether $B_\omega \equiv_m^P K^\omega$.

In Proposition 4.8 we have shown that

$$\text{SAT}^{\omega-ft} \leq_m^P \text{SAT}^{\omega-bf} \leq_m^P \text{SAT}^{\omega-t}.$$

It is not very likely that these sets are \equiv_m^P -equivalent. Thus we have the somewhat surprising suspicion that the complexity of the ω -jumps of the truth-table hierarchy depends heavily on the succinctness of the encodings of Boolean functions that we use. On the other hand, it turns out that $\text{SAT}^{\omega-\Delta} \equiv_m^P \text{SAT}^{\omega-bf}$, i. e. that the ω -jump of the difference hierarchy has the same complexity as one of the ω -jumps of the truth-table hierarchy.

THEOREM 5.1: (i) $\text{SAT}^{\omega-t}$ is $\leq_t^P(\text{NP})$ -complete.

(ii) $\text{SAT}^{\omega-bf}$ is $\leq_{bf}^P(\text{NP})$ -complete.

(iii) $\text{SAT}^{\omega-ft}$ is $\leq_{ft}^P(\text{NP})$ -complete.

Proof: As for Theorem 4.9 (i). ■

Let us emphasize once more that it is not likely that $\leq_{ft}^P(\text{NP}) = \leq_{bf}^P(\text{NP})$ or $\leq_{ft}^P(\text{NP}) = \leq_t^P(\text{NP})$ because these problems are closely related to the problems of whether there exist polynomial-time algorithms converting Boolean formulas or Boolean circuits into equivalent truth-tables. A similar dependency of the complexity of problems from the input presentation has been observed by other authors before. Sometimes an exponential increase of complexity can be observed if one chooses a succinct encoding of the instances of the problems (see [4] and [15]).

Next we show that $\text{SAT}^{\omega-\Delta}$ is also $\leq_{bf}^P(\text{NP})$ -complete. This is a consequence of a result in [16] where a sufficient condition is given for a set A to be complete in $\leq_{bf}^P(\text{NP})$. Note that this class is denoted there by $\mathbf{P}_{bf}^{\text{NP}}$.

THEOREM 5.2 [16]: *Let D be NP-complete and let $A \in \leq_{bf}^P(\mathbf{NP})$. If there exist a polynomial-time computable function f such that*

$$|\{i: x_i \in D\}| \text{ is odd} \leftrightarrow f(x_1, \dots, x_k) \in A$$

for all $k \geq 1, x_1, \dots, x_k \in \Sigma^$ such that $c_D(x_1) \geq c_D(x_2) \geq \dots \geq c_D(x_k)$ then A is $\leq_{bf}^P(\mathbf{NP})$ -complete. ■*

COROLLARY 5.3: $\text{SAT}^{\omega-\Delta}$ is $\leq_{bf}^P(\mathbf{NP})$ -complete. ■

In [16] many natural problems are shown to be $\leq_{bf}^P(\mathbf{NP})$ -complete, for example the problem $\{(G, n_1, \dots, n_k): k \geq 1 \text{ and the chromatic number of the graph } G \text{ is one of the numbers } n_1, \dots, n_k\}$.

It should be noted that our results do not depend particularly on NP and SAT as NP-complete problem. Just as well similar results can be developed for other classes having complete sets and being closed under union and intersection.

Finally let us mention some relationships between two classes investigated in [2] and the classes $\leq_{bf}^P(\mathbf{NP})$ and $\leq_{tt}^P(\mathbf{NP})$.

For a deterministic oracle Turing machine M with oracle set B and input x , let $Q(M, B, x)$ be the set of all queries which are asked by M to B during its computation on x . Further, let

$$Q(M, x) = \cup \{Q(M, B, x): B \text{ oracle set}\},$$

and

$$QA(M, x) = \cup \{Q(M, B, x): B \text{ oracle set and } M \text{ accepts } x \text{ with oracle } B\}.$$

For a set B , the set A is in $\mathbf{P.UNIV.ALL}(B)$ [$\mathbf{P.UNIV.ACC}(B)$] if and only if there exist two deterministic polynomial-time bounded Turing machines M and M' such that for every input x :

- $x \in A$ iff M accepts x using oracle B ,
- M' computes on input x the set $Q(M, x)$ [$QA(M, x)$].

Furthermore, define

$$\mathbf{P.UNIV.ALL} = \cup \{\mathbf{P.UNIV.ALL}(B): B \in \mathbf{NP}\}$$

and

$$\mathbf{P.UNIV.ACC} = \cup \{\mathbf{P.UNIV.ACC}(B): B \in \mathbf{NP}\}.$$

It follows from a result in [8] that $\mathbf{P.UNIV.ALL} = \leq_{tt}^P(\mathbf{NP})$. Moreover,

THEOREM 5.4 [2], [6]: For every set B ,

$$\mathbf{P.UNIV.ALL}(B) = \mathbf{P.UNIV.ACC}(B).$$

Proof: For the inclusion " \supseteq " see [2].

For the other inclusion we reproduce the proof from [6]. Let M_1 be a deterministic polynomial-time bounded oracle Turing machine accepting the set A with oracle B , and let M_2 be a deterministic polynomial-time bounded Turing machine computing on input x the set $Q(M_1, x)$. Now we choose an arbitrary word b from B , and we construct a deterministic polynomial-time bounded oracle Turing machine M'_1 such that

- M'_1 accepts the set A with oracle B ,
- $QA(M'_1, x) = Q(M_1, x) \cup \{b\}$ for all inputs x .

This implies $A \in \mathbf{P.UNIV.ACC}(B)$.

The Turing machine M'_1 works (with an arbitrary oracle C) as follows: On input x it computes first $Q(M_1, x)$ by simulating M_2 . Then it asks all $y \in Q(M_1, x) \cup \{b\}$ to the oracle C . Finally it simulates M_1 on input x . If x is accepted by M_1 then it is also accepted by M'_1 . If x is rejected by M_1 then x is rejected by M'_1 if and only if $b \in C$.

Evidently, for $C = B$ the machines M_1 and M'_1 accept the same set of words, namely A .

On the other hand, it is obvious that

$$QA(M'_1, x) \subseteq Q(M_1, x) \cup \{b\} \subseteq Q(M'_1, \Sigma^* - \{b\}, x) \subseteq QA(M'_1, x). \blacksquare$$

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