$P \neq NP$ for infinite time Turing machines

Ralf Schindler[a*]

[a] *Institut für formale Logik, Universität Wien, 1090 Wien, Austria*

rds@logic.univie.ac.at
http://www.logic.univie.ac.at/~rds/

Abstract. We state different versions of the $P =?NP$ problem for infinite time Turing machines. It is observed that $P \neq NP$ collapses to the fact that there are analytic sets which are not Borel.

In this note we study versions of the $P =?NP$ problem for infinite time Turing machines. The analytic sets of reals may be construed as an infinite analog to the class NP (cf. for example [3, §3.9]). The $P =?NP$ problems for infinite time Turing machines can therefore naturally be translated as questions about analytic sets. These questions have classical answers.

1 Analytic sets.

We shall have to consider Polish spaces, i.e., complete separable metric spaces. In particular we shall be interested in the Cantor space ${}^{\omega}2$ and in the Baire space ${}^{\omega}\omega$. We refer the reader to [2] for background information. However, the descriptive set theory which we shall need is pretty elementary indeed. In order to make the paper self-contained modulo [1] this section develops all the necessary descriptive set theoretic tools.

Let $\mathcal{X}$ be a Polish space and let $(\mathcal{O}_n : n < \omega)$ be a recursive enumeration of basic open sets. Let $\mathcal{O}^{\mathcal{X}} \subset \mathcal{X} \times {}^{\omega}2$ be defined by

$$(x, y) \in \mathcal{O}^{\mathcal{X}} \Leftrightarrow x \in \bigcup_{y(n)=1} \mathcal{O}_n.$$

*The author thanks J.D. Hamkins, R. Schipperus, and P. Welch for comments on an earlier version of this paper. He thanks Sanders Seliverstov for pointing out an inaccuracy in the first version of the proof of Theorem 2.10.
2000 *Mathematics Subject Classification.* 68Q15, 68Q17, 03E15.
Keywords: computer science/complexity theory/descriptive set theory.

Then $\mathcal{O}^{\mathcal{X}}$ is a universal open set, i.e., $\mathcal{O}^{\mathcal{X}}$ is open in $\mathcal{X} \times {}^{\omega}2$ and if $\mathcal{O}$ is an open subset of $\mathcal{X}$ then there is some $y \in {}^{\omega}2$ with

$$\mathcal{O} = \{x \mid (x, y) \in \mathcal{O}^{\mathcal{X}}\}.$$

A set $A \subset \mathcal{X}$ is analytic if and only if there is a closed $\mathcal{C} \subset \mathcal{X} \times {}^{\omega}\omega$ such that

$$x \in A \Leftrightarrow \exists y \in {}^{\omega}\omega \ (x, y) \in \mathcal{C}.$$

Let $\mathcal{U}^{\mathcal{X}} \subset \mathcal{X} \times {}^{\omega}2$ be defined by

$$(x, y) \in \mathcal{U}^{\mathcal{X}} \Leftrightarrow \exists z \in {}^{\omega}\omega \ ((x, z), y) \notin \mathcal{O}^{\mathcal{X} \times {}^{\omega}\omega}.$$

Then $\mathcal{U}^{\mathcal{X}}$ is a universal analytic set, i.e., $\mathcal{U}^{\mathcal{X}}$ is analytic in $\mathcal{X} \times {}^{\omega}2$ and if A is an analytic subset of $\mathcal{X}$ then there is some $y \in {}^{\omega}2$ with

$$A = \{x \mid (x, y) \in \mathcal{U}^{\mathcal{X}}\}.$$

Now let $\mathcal{X} = {}^{\omega}2$, and let us write $\mathcal{U}$ for $\mathcal{U}^{\mathcal{X}} = \mathcal{U}^{({}^{\omega}2)}$. The set $\Delta = \{x \mid (x, x) \in \mathcal{U}\}$ is analytic. If Δ were coanalytic, i.e., if ${}^{\omega}2 \setminus \Delta$ were analytic, then there would be some $y \in {}^{\omega}2$ with

$$\Delta = \{x \mid (x, y) \notin \mathcal{U}\};$$

but then $y \in \Delta$ iff $(y, y) \notin \mathcal{U}$ iff $y \notin \Delta$. Hence Δ is not coanalytic, and therefore not Borel.

Let $G \subset {}^{\omega}2$ be the set of all $x \in {}^{\omega}2$ which are not eventually constant (equivalently, such that there are arbitrary large m and m' with $x(m) = 0$ and $x(m') = 1$). G is a G_δ subset of ${}^{\omega}2$. We may define a bijection $\varphi \colon G \to {}^{\omega}\omega$ by $\varphi(x) = y$ if and only if the n^{th} block of 1's in x contains exacly $y(n)$ 1's. If $A \subset {}^{\omega}2 \times {}^{\omega}\omega$, say, then we may define $A^{\varphi} \subset {}^{\omega}2 \times {}^{\omega}2$ by

$$(x, y) \in A^{\varphi} \Leftrightarrow y \in G \wedge (x, \varphi(y)) \in A.$$

Now let $A \subset {}^{\omega}2$ be lightface analytic which means that there is a recursive $R(-, -)$ such that, if $\mathcal{C}$ denotes the closed set

$$[R] = \{(x, y) \in {}^{\omega}2 \times {}^{\omega}\omega \mid \forall n < \omega \ R(x \restriction n, y \restriction n)\},$$

then we have that

$$x \in A \Leftrightarrow \exists y \in {}^{\omega}\omega \ (x, y) \in \mathcal{C}.$$

We then also have that

$$x \in A \Leftrightarrow \exists y \in {}^{\omega}2 \ (x, y) \in \mathcal{C}^{\varphi}.$$

It is straightforward to verify that $\mathcal{C}^{\varphi}$ is no longer closed. Rather, $\mathcal{C}^{\varphi}$ is a lightface G_{δ} subset of ${}^{\omega}2$.

Now the set $\Delta \subset {}^{\omega}2$ defined above is a lightface analytic set. In fact, the witnessing recursive $R(-,-)$ is easily given by the complement of $\mathcal{O}^{{}^{\omega}2 \times {}^{\omega}\omega}$; we here need that $(\mathcal{O}_n : n < \omega)$ be recursive. We then have that

$$x \in \Delta \Leftrightarrow \exists y \in {}^{\omega}2 \ (x, y) \in \mathcal{G},$$

where $\mathcal{G}$ is a lightface G_{δ} subset of ${}^{\omega}2$. It is this latter representation of Δ which we shall need later on.

2 $P \neq NP$ for infinite time Turing machines.

Infinite time Turing machines were introduced in [1]. They have exactly the same hardware as traditional Turing machines. The difference is that one allows transfinite running times. We refer the reader to [1] for exact definitions. An acquaintance with §§1 and 2 (i.e., pp. 569-575) of [1] will basically suffice for our purposes. In what follows, by "Turing machine" we shall always mean an infinite time Turing machine.

It will be convenient to think of a Turing machine to come with *two* halting states, the accept state, and the reject state.

Definition 2.1 *Let $A \subset {}^{\omega}2$. We say that A is decidable in polynomial time, or $A \in P$, if there are a Turing machine T and some $m < \omega$ such that*
(a) T decides A (i.e, $x \in A$ iff T accepts x), and
(b) T halts on all inputs after $< \omega^m$ many steps.

With infinite time Turing machines, all inputs (i.e., elements of the Cantor space ${}^{\omega}2$) may be counted as having the same length, namely ω. So it appears reasonable to have a polynomial time Turing machine being one which always halts after $< \omega^m$ many steps, for some fixed $m < \omega$.

The following just generalizes Definition 2.1.

Definition 2.2 *Let $A \subset {}^{\omega}2$, and let $\alpha \leq \omega_1 + 1$. We say that A is in P_α if there are a Turing machine T and some $\beta < \alpha$ such that*
(a) T decides A (i.e, $x \in A$ iff T accepts x), and
(b) T halts on all inputs after $< \beta$ many steps.

Of course, $P = P_{\omega^\omega}$. Moreover, P_{ω_1+1} is just the class of all $A \subset {}^{\omega}2$ which are decided by some Turing machine.

Lemma 2.3 ([1, Theorem 2.6]) *Let $A \subset {}^{\omega}2$. Then $A \in P_{\omega^2}$ if and only if A is an arithmetic set.*

Lemma 2.4 *Let $A \subset {}^{\omega}2$. Then $A \in P_{\omega+2}$ if and only if A is a lightface G_δ set.*

Proof. This is straightforward. The less trivial direction is given by the proof of [1, Theorem 2.6]. □

Lemma 2.5 *Let $A \subset {}^{\omega}2$. Then $A \in P_{\omega_1^{CK}}$ if and only if A is a hyperarithmetic set. If $A \in P_{\omega_1}$ then A is a Borel set.*

Proof. The first part is [1, Theorem 2.7]. The second part is an immediate consequence of the proof thereof. □

It is on the other hand not true that every Borel set is in P_{ω_1}. Lemma 2.7 will characterize P_{ω_1}.

Definition 2.6 *Let $A \subset {}^{\omega}2$. If $\alpha < \omega_1$, then we say that $A \in \Delta^1_1(\alpha)$ if $A \in \Delta^1_1(x)$ uniformly for every real x coding α. We say that A is Δ^1_1 in a countable ordinal if there is some $\alpha < \omega_1$ such that $A \in \Delta^1_1(\alpha)$.*

Lemma 2.7 [1] *$A \in P_{\omega_1}$ if and only if A is Δ^1_1 in a countable ordinal. In fact, if α is admissible then $A \in P_\alpha$ if and only if $A \in \Delta^1_1(\beta)$ for some $\beta < \alpha$.*

Proof. This follows from revisiting the proof of [1, Theorem 2.7]. "$\Rightarrow$" is immediate. As to "$\Leftarrow$," note that we may pick a real x coding β such that $\alpha \geq \omega_1^x$ (= the least x-admissible $> \omega$). The Borel code for A is the a tree with rank $< \omega_1^x \leq \alpha$. □

We now turn to the class NP.

[1] This was observed independently by J.D. Hamkins.

Definition 2.8 *Let $A \subset {}^\omega 2$. We say that A is verifiable in polynomial time, or $A \in NP$, if there are a Turing machine T and some $m < \omega$ such that*
(a) $x \in A$ if and only if $(\exists y \;\; T$ accepts $x \oplus y)$, and
(b) T halts on all inputs after $< \omega^m$ many steps.

Definition 2.9 *Let $A \subset {}^\omega 2$, and let $\alpha \leq \omega_1 + 1$. We say that A is in NP_α, if there are a Turing machine T and some $\beta < \alpha$ such that*
(a) $x \in A$ if and only if $(\exists y \;\; T$ accepts $x \oplus y)$, and
(b) T halts on all inputs after $< \beta$ many steps.

Again, $NP = NP_{\omega^\omega}$. NP_α is the class of all projections of sets in P_α. It is now immediate that $P \neq NP$.

Theorem 2.10 $NP_{\omega+1} \setminus P_{\omega_2} \neq \emptyset$.

Proof. Let Δ and $\mathcal{G}$ be as in section 1. In particular, Δ is a lightface analytic subset of ${}^\omega 2$ which is not Borel, $\mathcal{G}$ is a lightface G_δ set, and

$$x \in \Delta \Leftrightarrow \exists y \in {}^\omega 2 \; (x, y) \in \mathcal{G}.$$

By Lemma 2.4, $\mathcal{G} \in P_{\omega+2}$. Hence $\Delta \in NP_{\omega+2}$. However, by Lemma 2.5, Δ cannot be in P_{ω_1}, as it is not Borel. □

Another version of the $P =?NP$ problem counts an input $x \in {}^\omega 2$ as having length ω_1^x (= the least x-admissible $> \omega$). Note that no admissible ordinal is clockable (cf. [1, Theorem 8.8]. This leads to:

Definition 2.11 *Let $A \subset {}^\omega 2$. We say that $A \in P^+$ if there is a Turing machine T such that*
(a) $x \in A$ if and only if T accepts x, and
(b) T halts on all inputs x after $< \omega_1^x$ many steps.

Definition 2.12 *Let $A \subset {}^\omega 2$. We say that $A \in NP^+$ if there is a Turing machine T such that*
(a) $x \in A$ if and only if $(\exists y \;\; T$ accepts $x \oplus y)$, and
(b) T halts on all inputs $x \oplus y$ after $< \omega_1^x$ many steps.

Again we'll have that $P \neq NP$.

Theorem 2.13 $P^+ = P_{\omega_1^{CK}} = \Delta^1_1$.

PROOF. Let $A \in P^+$. It is straightforward that there is then a Σ_1 formula Ψ (saying that there is a certain sequence of snapshots) such that

$$x \in A \Leftrightarrow L_{\omega_1^x}[x] \models \Psi(x).$$

This implies that A is coanalytic (i.e., Π^1_1). Of course, we also have that ${}^\omega 2 \setminus A \in P^+$, so that by the same argument ${}^\omega 2 \setminus A \in \Pi^1_1$. Therefore, $P^+ \subset \Delta^1_1$.

On the other hand, we have $\Delta^1_1 = P_{\omega_1^{CK}} \subset P^+$. $\square$

Corollary 2.14 *$NP^+ \setminus P^+ \neq \emptyset$.*

3 Some open problems.

We may allow a Turing machine to take even more time to reach its decision. Recall that if $\lambda + n$ is clockable for $n < \omega$ then so is λ. We arrive at:[2]

Definition 3.1 *Let $A \subset {}^\omega 2$. We say that $A \in P^{++}$ if there is a Turing machine T such that*
(a) $x \in A$ if and only if T accepts x, and
(b) T halts on all inputs x after $\leq \omega_1^x + \omega$ many steps.

Definition 3.2 *Let $A \subset {}^\omega 2$. We say that $A \in NP^{++}$ if there is a Turing machine T such that*
(a) $x \in A$ if and only if $(\exists y \;\; T$ accepts $x \oplus y)$, and
(b) T halts on all inputs $x \oplus y$ after $\leq \omega_1^x + \omega$ many steps.

P^{++} is a larger class than P^+:

Theorem 3.3 *Every lightface analytic set is in P^{++}.*

PROOF. Let A be a lightface analytic set. There is a recursive $R(-,-)$ such that

$$x \in A \Leftrightarrow \exists y \in {}^\omega\omega \; \forall n < \omega \; R(x \restriction n, y \restriction n).$$

[2] In spirit this has been suggested by P. Welch.

For $x \in {}^\omega 2$ consider the tree

$$T_x = \{s \mid R(x \upharpoonright lh(s), s)\}.$$

Then

$$x \in A \Leftrightarrow T_x \text{ is illfounded.}$$

We can design a Turing machine $\mathcal{T}$ which, on input x, first produces T_x and then crosses out the wellfounded part of T_x. This wellfounded part has rank $\leq \omega_1^x$ (as every wellfounded tree which is recursive in x has rank $< \omega_1^x$). The machine $\mathcal{T}$ is finally suppose to check if there is something left after crossing out the wellfounded part of T_x. This will take another ω many steps of computation. On input x, $\mathcal{T}$ has therefore a running time $\leq \omega_1^x + \omega$. $\square$

Of course, P^{++} is also closed under complements.

Question. $P^{++} \neq NP^{++}$?

Definition 3.4 *Let $f\colon \mathcal{D} \to \omega_1$. Let $A \subset {}^\omega 2$. We say that $A \in P^f$ if there is a Turing machine T such that*
(a) $x \in A$ if and only if T accepts x, and
(b) T halts on all inputs x after $< f(x)$ many steps.

Definition 3.5 *Let $f\colon \mathcal{D} \to \omega_1$. Let $A \subset {}^\omega 2$. We say that $A \in NP^f$ if there is a Turing machine T such that*
(a) $x \in A$ if and only if ($\exists y$ T accepts $x \oplus y$), and
(b) T halts on all inputs $x \oplus y$ after $< f(x)$ many steps.

Question. For which $f\colon \mathcal{D} \to \omega_1$ is $P^f \neq NP^f$?

References

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